RUNTIME TECHNIQUES AND INTERPROCEDURAL ANALYSIS IN JAVA VIRTUAL MACHINES

by

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Abstract

Java programs are deployed in a bytecode format that is executed by a Java Virtual Machine (JVM). JVM performance is determined by several major components: execution engine, garbage collector, and threading system. In this thesis, we study a new garbage collection scheme and speculative optimizations in just-in-time (JIT) compilers for improving Java performance.

We present a novel approach for reducing garbage collection frequencies. Without an escape analysis, the system uses write barriers to capture escaping objects and allocation sites. Instead of allocating non-escaping objects on stacks, the system allocates them in regions that are treated as extensions of stack frames. By freeing regions with associated stack frames, the system can reduce the frequency of garbage collections. We present the overall idea and provide details of a specific design and implementation.

A JVM allows dynamic class loading. A JIT compiler can speculatively optimize code base on loaded classes only. However, the virtual machine must revert speculatively optimized code if newly loaded classes invalidates optimization assumptions. In this thesis, we review existing techniques supporting speculative optimizations, including runtime guards, code patching, and on-stack replacement. We present an improvement and implementation of an on-stack replacement mechanism.

A call graph is necessary for developing interprocedural program analyses. Call graph construction in a Java virtual machine needs to overcome difficulties of dynamic class loading and lazy reference resolution. In this thesis, we show a general approach to adapt static type analysis to a JIT environment for building call graphs. We also introduce a new call graph profiling mechanism using code stubs.

We developed a general type analysis framework for supporting *speculative method inlining* in a JIT environment. Several popular type analyses were implemented in the framework, including two interprocedural ones, XTA and VTA. Using the framework, we did an extensive limit study of method inlining and reported our findings and experience in the thesis.

In each chapter we discuss the related work. At the end of the thesis, we point out future research directions.

Résumé

French abstract goes here.

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Chapter 1
Introduction

The Java programming language gains popularity in many application domains. Java's execution model is different from many static compiled languages such as Fortran and C/C++. The Just-In-Time (JIT) compilation approach deployed by Java virtual machines (JVMs) creates a new scenario for language implementation and compiler optimizations.

In this thesis, we study both challenges and opportunities of program analyses and compiler optimizations in Java virtual machines. The main contents of this thesis include a new garbage collection scheme using dynamic techniques and interprocedural program analyses in the presence of dynamic class loading and lazy method compilation. Dynamic analyses are designed to support speculative optimizations for best performance results. The problems and solutions can also be applied to virtual machines for other languages such as `C#`.

1.1 The Java programming langauge and the Java virtual machine

Since it was born as a programming language for small devices, the Java programming language [AGH00] has become a mainstream programming language in many domains, such as web and e-business application development, transaction server implementation, scientific computation, etc. Java[1] is designed as a high-level object-oriented language with safety features. Type checking is required at both compilation time and execution time. Null pointer accesses and array bounds overflows must be captured and reported to applications. Garbage collection is the default memory management, which frees programmers from freeing objects manually and reduces the chances of human mistakes. Other desirable features include multi-threading and large common libraries.

The Java virtual machine (JVM) specification [LY96] is designed to support the Java programming language, but it does not interact with the source language directly. As indicated by its name, the Java virtual machine is an abstract machine which defines an instruction set, bytecodes, and other resources.

It is necessary to introduce the structure of the Java virtual machine before examining the technical details of its implementation. A Java thread is conceptually similar to an operating system thread. Execution in a thread is sequential, but multiple threads can execute in parallel. Each thread has its own program counter and private stack. The *heap* is a shared data area among all Java threads. Objects and arrays are allocated in the heap and reclaimed by garbage collectors.

The implementation of the Java virtual machine specification is architecture-dependent. It can be a piece of software [Gri98, jikb, sab, kaf] written in other languages, such as assembly, C, or even Java itself, or implemented as hardware CPUs [aji, jav]. The software approach is more prominent in practice. A software virtual machine contains a runtime system that supports threading, garbage collection, networking, IO support, etc. The bytecode instructions can be interpreted. An alternative is to compile the bytecode to native code and execute the native code. Often the latter approach is called Just-In-Time (JIT) compilation. Note that the JVM specification supports classes from languages other than Java if the classes obey the specification.

Java applications are compiled to and distributed in a compact, platform-neutral representation, the `class` file format. The `class` file format is the binary format

[1]We refer to the Java programming language as `Java`, and the Java virtual machine as `JVM` in this thesis.

accepted by Java virtual machines. Each class file has a constant pool which contains symbolic names of types, methods, and fields accessed by bytecode instructions of methods in this class. Bytecode instructions are stack-based. We use a simple example, `Helloworld.java`, to facilitate the understanding of the class file format.

```
1: public class Helloworld {
2:   public static void main(String[] args) {
3:     System.out.println(``Hello world!'');
4:   }
5: }
```

The `Helloworld.java` is compiled to a binary class file "`Helloworld.class`" by a Java-to-bytecode compiler such as `javac`, `jikes` [jika], or `soot` [soo]. Figure 1.1 are partial contents of the class file, disassembled by the command line "`javap -v Helloworld`". At the bytecode index 5 of the `main` method is an "`invokevirtual #4`" instruction, which corresponds to the invocation of "`println`" method at line 3 in the source code. Note that the number "`#4`" in the bytecode instruction is an index of the constant pool of `Helloworld.class`. Further, the `#4` entry of the constant pool is comprised by entries `#19` and `#20`. Recursively tracing down, the `#4` entry represents a method signature "`java/io/PrintStream.println(Ljava/lang/String;)V`".

From above example, we see that bytecode instructions access types, methods and fields by their symbolic names. The virtual machine is responsible for resolving these names to concrete classes, methods or fields at runtime.

The Java virtual machine supports two kinds of types: *primitive types* and *reference types*. Primitive types include integral types and float types. Reference type variables are pointers to objects or arrays. A bytecode instruction explicitly distinguishes the types of values it operates on. For example, an *aload* instruction copies an object reference from a variable to the stack, and an integer value is loaded by an *iload* instruction. Although the class file does not contain declaring types of local variables, the virtual machine can reconstruct static types for variables using the type information contained in bytecode instructions. Rich type information in bytecodes is not only necessary for type checking, but also very useful for efficient garbage collections and compiler optimizations.

```
public class Helloworld extends java.lang.Object
  Constant pool:

const #2 = Field        #16.#17;
const #3 = String       #18;
const #4 = Method       #19.#20;

const #16 = class       #23;
const #17 = NameAndType #24:#25;
const #18 = Asciz       Hello world!;
const #19 = class       #26;
const #20 = NameAndType #27:#28;

const #23 = Asciz       java/lang/System;
const #24 = Asciz       out;
const #25 = Asciz       Ljava/io/PrintStream;;
const #26 = Asciz       java/io/PrintStream;
const #27 = Asciz       println;
const #28 = Asciz       (Ljava/lang/String;)V;

public static void main(java.lang.String[]);
  Code:
   Stack=2, Locals=1, Args_size=1
   0:   getstatic       #2;
   3:   ldc             #3;
   5:   invokevirtual   #4;
   8:   return
  LineNumberTable:
   line 3: 0
   line 4: 8
```

Figure 1.1: Partial contents of `Helloworld.class`

The execution of a Java application starts at a special method `public void main(String[])` of a class whose name is supplied by the command line. Before using a class (e.g., creating instances, accessing members, etc.), the virtual machine must load the class and create an internal representation for it. The process involves several steps:

Loading. The class loading process finds the binary representation of the class. The binary can come from various sources, depending on the class loader used. Then the virtual machine parses the representation according to the `class` file format. This process also checks some constraints required by the virtual machine specification. If the class can be loaded, and it has a direct superclass or super interface, the symbolic references to them must be resolved. This resolution triggers the loading of the superclass or super interface if necessary.

Linking. The linking process verifies the representation is well formed. The virtual machine may use some data-flow analyses to verify that bytecode instructions satisfy some semantic requirements. Then the virtual machine creates necessary internal data structures, such as type information and method tables, as the internal representation of the class. There are several kinds of symbolic references in the constant pool. In particular, we are interested in type, method, and field references. Resolving a symbolic reference may cause other classes to be loaded. The Java virtual machine specification does not require symbolic references to be resolved at linking time. Instead, many VM implementations choose the laziest strategy for saving resources and improving responsiveness. By the laziest strategy, symbolic references are not resolved until used. Thus, dynamic class loading can be triggered by the execution of instructions for object creation, method invocation, field accesses and type checking.

Initialization. The initialization process executes the static initializer method "`<clinit>`" and initializes static fields of the class. It may trigger the initialization of its superclass and super interfaces.

A class may contain several methods. An *abstract* method has only a name, parameter and return types. A *normal* method has an array of bytecode instructions. The invocation of a normal method requires its bytecode instructions to be parsed or compiled first. The compilation of bytecode instructions does not need to be done when loading and resolving the class. Instead, to save the workload and resources, a virtual machine can delay the process until the first invocation of the method during the execution. This technique is typically called *lazy method compilation*.

The Java programming language provides constructs for creating objects in heaps, but there are no constructs for deallocating objects. Heaps are managed by automatic memory managers, such as garbage collectors. The Java virtual machine specification has no specific requirements on choosing which garbage collection algorithm to use.

1.2 Motivation

Two features distinguish Java from other languages: automatic memory management (garbage collection) and Just-In-Time (JIT) compilation. Program analyses have been used in nearly every compiler for optimizations. Previous research on program analysis mainly focused on ahead-of-time compilers. JIT compilation mode presents a different scenario for program analysis designers. In this thesis, we refer to a program analysis performed in an ahead-of-time compiler as *static*, and the one in a JIT compiler as *dynamic*.

1.2.1 Automatic memory management

The Java virtual machine implicitly uses garbage collectors as default memory managers. Garbage collection [JL96] has the advantage of freeing memory safely and often precisely. However, tracing and collecting objects consumes many machine cycles during the program execution, so reducing the frequency of garbage collection can be beneficial.

Stack-allocation of objects is one promising approach to reduce the work of a garbage collector [GS00]. If an object does not escape a method, it can be created on a stack frame instead of in a heap. Objects on the stack can be reclaimed without intervention by the garbage collector. However, there are several potential obstacles to performing object stack-allocation in the Java virtual machine: existing techniques require static escape information, there are restrictions on objects that can be allocated on stack, and *finalize* methods have to be handled specially. This motivates us to design a new object allocator reducing garbage collection overhead without above restrictions in Java virtual machines.

In our initial design, we use runtime checks to detect escaping objects. Heap organization must be changed to facilitate new allocation scheme. This approach allows us to study benchmark behaviors using the new allocator. It also gives us an indication how to design a runtime escape analysis for eliminating unnecessary checks.

1.2.2 JIT compilation and optimizations

Over the last 10 years, virtual machine technology has greatly advanced. Most Java virtual machines running on desktop or server computers have built-in Just-in-time compilers, which obtain 10 fold or more performance improvement than an interpreter. Most *intra*procedural data-flow analyses have been implemented in these JIT compilers [PVC01, AAB$^+$00, SOT$^+$00, CLS00, GKS$^+$04]. Further performance improvements have been achieved using adaptive and feedback-directed optimizations [AFG$^+$00, AHR02]. To improve startup performance, a mixed mode of interpretation and JIT compilation is often used to interpret a method first and only selectively compile hot methods [PVC01, AAB$^+$00].

Methods are commonly used as compilation units. Not only do methods provide a natural semantic boundary of code blocks, but also the implementation of dynamic compilation becomes easier in the presence of virtual method calls. State-of-the-art JIT compilers have implemented most intraprocedural (within a method) analyses and optimizations as in ahead-of-time (static) compilers. Dynamic *inter*procedural (over a collection of methods) analyses, however, have not yet been widely adopted. Very few coarse type-based interprocedural analyses (IPAs) [IKY$^+$00, PS01] have gained ground in JIT compilation environments. However, work relating to more complicated, reachability-based IPAs, such as dynamic points-to analysis and escape analysis, is only just starting to emerge [HDH04, QH04, QH05].

Static IPAs assume that the whole program is available at analysis time. However, this may not be the case in a Java virtual machine. The virtual machine can download a class file from the network, create it on-the-fly, or retrieve it from other unknown resources. Even when all programs exist on local disks, the virtual machine typically loads classes lazily, on demand, to reduce resource usage and improve responsiveness [LB98]. When a JIT compiler encounters an unresolved symbolic reference, it may choose to delay resolution until the instruction is executed at runtime. These unresolved references has to be dealt with correctly by dynamic IPAs.

Although Java's dynamic features pose difficulties for program analyses and optimizations, there are many opportunities at runtime that can only be enjoyed by dynamic analyses. For example, a dynamic IPA only needs to analyze loaded classes and invoked methods. Therefore, the analysis can be more efficient and the results are more precise comparing to a static analysis which must analyze programs based on conservative assumptions about which classes might be loaded and which methods might be invoked. Thus, a dynamic IPA's analyzed code base can be much smaller than in a conservative static analysis. Further, dynamic class loading can improve the precision of type analyses. The set of runtime types can be limited to loaded classes. Thus, a dynamic analysis has more precise type information than its static counterpart. In contrast to the conservative (pessimistic) nature of static analysis, a dynamic one can be optimistic about future execution, if used in conjunction with runtime invalidation mechanisms [PS01, IKY^{+}00, FQ03, SYN03a].

Dynamic IPAs seem more suitable for long-running applications in adaptive recompilation systems. Pechtchanski and Sarkar [PS01] described a general approach of using dynamic IPAs. A virtual machine gathers information about compiled methods and loaded classes in the initial state, and performs recompilation and optimizations only on selected hot methods. When the application reaches a "stable state", information changes should be rare.

In summary, the development of more advanced interprocedural analyses in JIT environments has not been widely explored and practiced. The differences between static and dynamic interprocedural analyses are:

1. a static analysis has the full program available to the analysis whereas a dynamic one only has seen the executed part;

2. a dynamic analysis has much tighter limitations on space and time resources, but a static analysis is more flexible in general; and

3. a static analysis has to be conservative, but a dynamic one can be speculatively optimistic if the system has the ability to invalidate the code or execution states when optimistic assumptions are violated.

One goal of this thesis is to design advanced dynamic interprocedural analysis at runtime for improving Java performance in the full context of the Java virtual machine specification. We also study techniques for supporting speculative optimizations using interprocedural analysis results.

1.3 Contributions

This thesis made following contributions to the virtual machine research area:

Region-based allocator. Without an effective online escape analysis, the effect of object stack-allocation is limited in Java virtual machines. Instead, we suggested an adaptive region-based allocator in Chapter 3. Our approach uses dynamic write barriers to detect escaping objects. By extending a stack frame with a region, other restrictions of object stack-allocation are removed. We had implemented an prototype in an early version of Jikes RVM, and we studied detailed behaviors of the allocator.

Improvement and implementation of an on-stack replacement algorithm. Speculative optimizations may yield better performance improvement than conservative optimizations. To support speculative optimizations, a Java virtual machine needs invalidation mechanisms as backups for correcting wrong speculations. In Chapter 4, we reviewed several existing invalidation techniques and presented an improvement and implementation of a new on-stack-replacement mechanism [FW00] in Jikes RVM.

Efficient call graph construction in the presence of dynamic class loading and lazy compilation. Interprocedural analysis needs a call graph of the program. In Chapter 5, we studied call graph constructions in Java virtual machines. We demonstrated a general approach for handling dynamic class loading in a dynamic program analysis. We did a detailed comparison study of several well-known type analyses for constructing call graphs. Furthermore, we designed and evaluated a novel mechanism [QH04] for constructing accurate call graphs with cheap cost. All mechanisms have been implemented in Jikes RVM and evaluated on a set of standard Java benchmarks.

Dynamic interprocedural type analyses and method inlining. We conducted a thorough study of speculative method inlining in Chapter 6. First, we presented a limit study of method inlining using type analyses. We analyzed the strength and weakness of each analysis. Using runtime call graphs, we developed two advanced interprocedural type analyses in a JIT environment. We showed an incremental, event-driven model of dynamic interprocedural analysis which handles dynamic class loading and lazy reference resolution properly. We also pointed out strength and weakness of simple class hierarchy analysis and dynamic interprocedural type analysis.

1.4 Thesis outline

First, in Chapter 2, we briefly introduce Jikes RVM, the test bed of our implementations and benchmarks used in this thesis. Chapter 3 introduces the design and evaluation of a region-based allocator. We review runtime techniques supporting speculative inlining in Chapter 4. Improvement and implementation of an on-stack replacement algorithm is presented in this chapter as well. We study different dynamic call graph construction algorithms in Chapter 5, which includes several dynamic type analyses and a new profiling mechanism. Chapter 6 studies method inlining using type analysis. This chapter also presents the design and evaluation of a reachability-based interprocedural type analysis as an application of dynamic call graphs. Finally conclusions and future work are given in Chapter 7.

Chapter 2
Setting

A programming language requires an efficient implementation to prove it is useful. Java's inventor, Sun Microsystem, encourages the model of open design and private implementation of the Java virtual machine specification [LY96]. There is a variety of Java implementations accessible. We chose an open source virtual machine, Jikes RVM [jikb], as our test bed for its maturity and active support. The Jikes RVM implementation includes an efficient runtime system, a simple baseline compiler and an advanced optimizing compiler, a collection of GC implementations, and rich documentation.

2.1 Jikes Research Virtual Machine (RVM)

Jikes RVM [AAB+00, jikb] is an open-source (under IBM's Public Licence [CPL]) research virtual machine for executing Java bytecodes. Jikes RVM implements most of the Java virtual machine specification, and can run a variety of Java benchmarks. Jikes RVM itself is mostly written in Java, including compilers, runtime system, and garbage collectors. RVM classes are in a special package `com.ibm.JikesRVM`. Jikes RVM uses a public class library `GNU classpath` [cla], which is independent of virtual machines (a virtual machine needs to provide a small number of proxy classes in order to use the library). The bootstrap code and low-level signal handling functions are written in C. Jikes RVM currently supports four OS/architectures: AIX/PowerPC, Linux/x86, Linux/PowerPC, and Mac OS X. Because of its openness, maturity, and active support, Jikes RVM is an ideal test bed for experimenting new VM technologies.

2.1.1 Compilers

Jikes RVM provides two compilers, a *baseline* compiler and an *optimizing* compiler. It uses a *compile-only* approach where bytecode instructions are compiled to native code before execution. The baseline compiler has some common aspects to a bytecode interpreter: fast compilation and low performance. It generates machine code quickly in a single pass, but the code quality is relatively poor. In fact, the baseline compiled code simulates the stack architecture outlined in the specification [LY96].

The optimizing compiler performs both static data-flow analyses and feed-back directed optimizations on compiled methods. Optimized code is as efficient as those produced by industrial JIT compilers such as Sun's HotSpot server compiler [PVC01] and IBM's product JIT compiler [SOT$^+$00].

The bootstrapping process of Jikes RVM includes compiling RVM source code (in Java) to standard Java class files using the `jikes` [1] compiler. A bootimage is a binary executable which contains the baseline compiler, a system class loader, garbage collectors, and optionally other components. Chosen components build a list of class files to be initialized and compiled by a bootimage writer tool (written in Java) on a host Java virtual machine. Objects on the host VM are converted to RVM objects and methods in the bootimage classes are pre-compiled to machine code by RVM compilers. RVM classes, objects, and machine code are written into the bootimage. With some C and assembly code, the bootimage writer builds a binary executable bootimage of Jikes RVM which can run Java applications as other virtual machines.

When executing an application, the bootimage loads itself entirely into the heap and executes some initialization code. Then it parses the command line and finds the main class of the application. A main thread is created for the application and `public static void main(String[])` method of the main class is invoked. The bootimage also creates several system threads for garbage collectors, recompilation

[1]Jikes [jika] is a Java-to-bytecode compiler, and Jikes RVM is a Java virtual machine.

system, adaptive controller, debugger, etc. After the main thread terminates, RVM unloads resources and shuts down itself.

The Jikes RVM web site [jikb] has considerable information about the design and implementation of the system. In this part of the thesis, we introduce some technologies used in the virtual machine, which are highly related to our thesis contents.

Internal representations. It is necessary to understand the object layout and the internal representation of classes in Jikes RVM. We use the example in Figure 2.1 to explain the idea. A resolved class `A` has a *TypeInformationBlock* (TIB), which contains superclass and super interface ids, interface method table, virtual method table, and other miscellaneous information. There is a global data structure called Java Table Of Contents (JTOC), whose entries are values of literals, static fields, or machine code addresses of static methods. The start address of JTOC is kept in a register for fast access. Each class member (field or method) is assigned a runtime constant `offset` at class resolution time. The `offsets` of static members are used to access contents in JTOC (dashed lines in Figure 2.1). For example, `getstatic A.s_f` is simply compiled to instructions:

```
v = *(JTOC + A.s_f's offset);
```

The offset of a non-static field, e.g, `A.i_f`, is used to calculate the address of the field value when given an object pointer of type `A`. The offset of virtual method (`A.v_m`) decides where its machine code address locates in `A`'s virtual method table.

Lazy compilation. To reduce compilation overhead, Jikes RVM delays the compilation of a method until its first time invocation. This is done by, at class resolution time, putting the address of a special code stub, *trampoline*, in the virtual method table or JTOC instead of eagerly compiling the method. When the method is invoked, the special code stub is executed. The code stub blocks the execution and calls the compiler to compile the method. After the compilation was done, the virtual machine fills the method's entry in virtual method table or JTOC with its machine code address. Then the code stub resumes execution by jumping to the compiled code directly. Since the method's entry in virtual method table or JTOC has been replaced by real machine code address, future invocation on the method directly jumps to the machine code address.

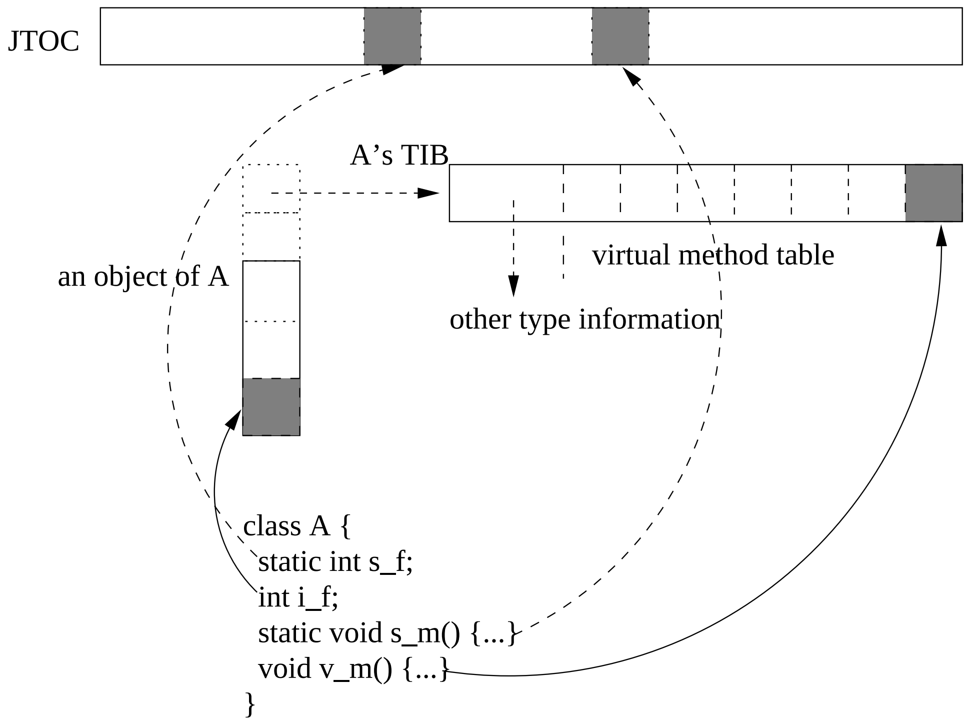


Figure 2.1: Class and object layout in Jikes RVM

Dynamic linking. As we discussed in Section 1.1, the class file contains symbolic references of types, methods and fields used by bytecode instructions. A reference must be resolved to a concrete entity before executing the instruction. Resolving a reference may trigger the loading of other classes. A virtual machine can take the laziest strategy to delay the resolution until the instruction gets executed. Jikes RVM takes such a lazy approach, called dynamic linking. Efficient implementation of dynamic linking requires cooperation between compilers and the runtime system.

When compiling a bytecode instruction accessing a field or method reference (e.g., `getfield`, `invokevirtual`, etc.), the compilers checks if the reference can be resolved without loading other classes. A resolved field or method has an `offset` for accessing its value or machine code address. Therefore, the compilers can generate efficient instructions to access a resolved member's value by its offset.

Each field or method reference is assigned a unique id when it is created, and the virtual machine maintains a table of offsets for unresolved ones. The table is indexed by the unique id of each reference, and the contents are initialized to a special value. When a reference gets resolved, the table entry is set to the resolved entity's offset. When compiling a bytecode instruction accessing a reference that cannot be resolved at compile time, the compilers generate instructions for checking if the table entry of the reference contains a valid offset value. If not, there is an instruction calling a resolution method to resolve the reference. Otherwise, the offset is read out from the table and used to access the member.

Compiler IRs. The baseline compiler does one-pass parsing of bytecodes and generates machine code quickly. The compiler has a big loop and code generation mimics the stack architecture defined by the Java virtual machine specification. Although the baseline compiler is easy to understand and modify, the stack nature of bytecode makes the conventional data-flow analysis harder. On the other hand, the optimizing compiler compiles bytecodes to machine code through several intermediate representations (IRs). It performs many data-flow analyses and optimizations on each IR. The IRs include high-level IR (HIR), low-level IR (LIR), and machine code level IR (MIR). Optionally there is a Static Single-Assignment (SSA) [Muc97] form available at HIR and LIR. Developing data-flow analyses based on HIR or LIR is much easier than raw bytecodes.

Compiler optimizations. The optimizing compiler provides a full suite of optimizations that include standard data-flow optimizations such as constant propagation and folding, dead code elimination, etc. There are some other Java-specific optimizations such as null check and bounds check eliminations. Method inlining is an important optimization for object-oriented programs. The optimizing compiler performs both static inlining (using type analyses results) and adaptive inlining (using profiling information). The default type analysis used for static inlining is the class hierarchy analysis (CHA) [DGC95]. It also implements method and class tests [DA99], code patching [IKY+00], pre-existence based inlining [DA99].

2.1.2 Experimental approach

The product configuration of Jikes RVM compiles all RVM classes into the bootimage. The execution uses a mixed compilation mode. Methods are quickly compiled by the baseline compiler first. Only hot methods are selected for recompilation by the optimizing compiler. Jikes RVM uses an adaptive analytical model [AFG+00] for driving recompilation and optimizations. The estimation of costs and benefits is based on samples collected at thread-switch points. The thread-switching is driven by OS timers. Therefore, the behaviour of the adaptive system is subject to OS workload and is nondeterministic. Although the default adaptive analytic model is flexible and intelligent, the decision can be easily affected by small changes made in the virtual machine code. For example, in one of our early experiments, the change we made in the baseline compiler slowed down the application at startup time. Without retraining the adaptive system, the recompilation decision was changed dramatically.

In order to compare results quantatively, we used a counter-based strategy for recompilation in our experiments in this thesis. Hot methods are selected for recompilation based on their invocation counters. Thus, nondeterminism introduced by the virtual machine is mostly eliminated. We verified that the recompilation behaviors between different runs of the same benchmark are very similar.

2.2 Benchmarks

In this section, we introduce the benchmark suite we used in our experiments. SpecJVM98 [speb] is a client-side benchmark suite for experimenting Java virtual machine development. It consists of 8 benchmarks introduced in Table 2.1[2]. The suite also provides a driver class, `SpecApplication` to execute individual benchmark with a number of iterations without restarting the virtual machine between runs. This can be used to simulate a long running application. Usually the first a few runs let the virtual machine compile most of the executed methods, recompile and optimize a few hot methods. After a few runs, there are less VM activities. The performance of later runs can be used to measure the quality of optimized code. Appendix A has a summary of key metrics of benchmarks [3].

SpecJBB2000 [spea] is a server-side benchmark which emulates a 3-tier system with emphasis on the middle tier. It models a wholesale company and supports several warehouses. Several clients send operation requests to the server and each client operates on a dedicated warehouse. The server creates one thread for each client. All warehouse data are resident in the heap. SpecJBB2000 is a multi-threading, long-running Java server benchmark.

In addition to the standard Spec benchmark suites, we used several benchmarks in different experiments. `Soot-c` [VRGH+00] is a Java bytecode transformation framework that is quite object-oriented, and which has several phases with potentially different allocation behaviors. `CFS` is a correlation-based feature subset selection evaluator from a popular open-source data mining package Weka [wek]. The program has an object-oriented design and does intensive numerical computation. We use a driver similar to the one from SpecJVM98 to run the CFS several times. The first run reads the data from a file and following runs operate the data on the heap. `Simulator` [cer] is a certificate revocation schemes. A variation of simulator interwoven with AspectJ [aspa] code is also used in some of our experiments.

[2]The description comes from `http://www.spec.org`.

[3]For more metrics, see `http://www.sable.mcgill.ca/metrics/`

_201_compress	Modified Lempel-Ziv method (LZW). It finds common substrings and replaces them with a variable size code.
_202_jess	JESS is a Java Expert Shell System. The benchmark workload solves a set of puzzles.
_205_raytrace	A raytracer that works on a scene depicting a dinosaur.
_209_db	Performs multiple database functions on memory resident database. Reads in a 1 MB file which contains records with names, addresses and phone numbers of entities and a 19KB file called scr6 which contains a stream of operations to perform on the records in the file.
_213_javac	This is the Java compiler from the JDK 1.0.2.
_222_mpegaudio	This is an application that decompresses audio files that conform to the ISO MPEG Layer-3 audio specification.
_227_mtrt	This is a variant of _205_raytrace, where two threads each renders the scene in the input file time-test model, which is 340KB in size.
_228_jack	A Java parser generator that is based on the Purdue Compiler Construction Tool Set (PCCTS). This is an early version of what is now called JavaCC.

Table 2.1: Introduction of SpecJVM98 benchmarks

Chapter 3
Region-Based Allocator

This chapter introduces an adaptive, region-based allocator for Java. The basic idea is to allocate non-escaping objects in local regions, which are allocated and freed in conjunction with their associated stack frames. By releasing memory associated with these stack frames, the burden on the garbage collector is reduced, possibly resulting in fewer collections.

The novelty of our approach is that it does not require static escape analysis, programmer annotations, or special type systems. The approach is transparent to the Java programmer and relatively simple to add to an existing JVM. The system starts by assuming that all allocated objects are local to their stack region, and then catches escaping objects via write barriers. When an object is caught escaping, its associated allocation site is marked as a non-local site, so that subsequent allocations will be put directly in the global region. Thus, as execution proceeds, only those allocation sites that are likely to produce non-escaping objects are allocated to their local stack region.

We present the overall idea, and then provide details of a specific design and implementation in Jikes RVM. Our experimental study evaluates the idea using the SpecJVM98 benchmarks, plus one other large benchmark. We show that a region-based allocator is a reasonable choice, that overheads can be kept low, and that the adaptive system is successful at finding local regions that contain no escaping objects.

3.1 Overview

The whole system consists of three parts: the allocator manages regions and allocates space for objects; the JIT compiler inserts instructions for acquiring and releasing a region in each compiled method; and the collector performs garbage collection when no more heap space is available.

In a region system, the heap space is divided into pages. The pages can be fixed-size or variable-size. In our system, we use fixed-size pages for fast computation of page numbers from addresses. The allocator is also a region manager. It manages a limited number of tokens. Each token is a small integer number identifying a region. Two regions, GLOBAL and FREE, exist throughout the execution of a program. Other, local regions, exist for shorter durations. They are assigned to and released by methods dynamically.

A high-level view of our memory organization is given in Figure 3.1. A more detailed description of the implementation is given in Section 3.2.

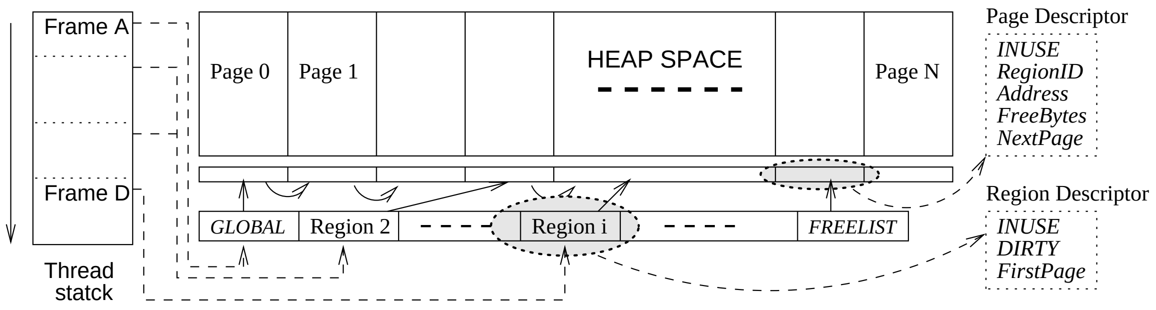


Figure 3.1: Memory organization of page-based heaps with regions

A region token points to a list of pages in the heap. The region space is expanded by inserting free pages at the head of the list. The GLOBAL region contains objects created by non-local allocation sites and pages containing objects that have escaped out of local regions. The GLOBAL region space can only be reclaimed by the collector. The system uses bit maps to keep track of free pages in the heap. The pages of a local region can be appended to the GLOBAL region or reclaimed by resetting their entries in the bit map.

A method activation obtains a region token by either acquiring an available token from the region manager or by inheriting its caller's token. The region identified by the token acts as an extension of the activation's stack frame. Before exiting, the activation releases the region if it was not inherited. It is clear that the lifetime of a local region is bounded by the lifetime of its host stack frame. There is a many-to-one mapping between stack frames and regions.

An object can be created in the region of the top stack frame or in the GLOBAL region. For the remainder of the discussion we need to define what we mean by an object *escaping* from a region, and a *non-local* allocation site.

Definition 1 *An object* escapes its allocation region *if and only if it becomes pointed to by an object in another region.*

Definition 2 *An allocation site becomes* non-local *when an object created by that site escapes.*

Given this definition of escape, there are only three Java bytecode instructions, `putstatic`, `putfield`, and `aastore`, that can lead to an object escaping. Therefore, it is sufficient to insert write barriers before these instructions to detect the escape of an object.

There is one additional situation that must be considered. When a method returns an object, the object may escape its allocation region via stack frames. However, this kind of escape can be prevented by either: (1) inserting write barriers before all `areturn` byte codes, or (2) requiring all methods returning objects to inherit their caller's region. In our implementation we have taken the second approach. It should be noted that objects passed to the callee as parameters are not a problem since the lifetime of the callee's stack frame is bounded by the caller's.

For an assignment such as $lhs.f = rhs$, the write barrier checks if the rhs is in the same region as the lhs object. When they are in different regions, the region containing the rhs object is marked as dirty. Since static fields are much like global variables, we assume that a *putstatic* always leads to the rhs object escaping, and the region associated with this object is marked as dirty.

It is worth pointing out that a region cannot contain an object reachable from other regions without being marked as dirty. If there is a path which causes an object o_1 of a region R_1 to be reached from objects in other regions, there must be an object, say o_i, in R_1 which is on the path and is directly pointed to by another object not in R_1, and the assignment of this pointer must have been captured by write barriers. Hence, R_1 must be marked as dirty when such a path exists.

Each allocation site in a compiled method is uniquely indexed, and each object has a field in its header for recording the index of its allocation site (see Section 3.3 for a discussion of how this is accomplished without increasing the object header size). The allocator maintains a bit vector to record the states of the allocation sites. Besides marking the region dirty, the write barrier also marks an escaping object's allocation site as *non-local*. The allocator allocates objects in local regions only for local allocation sites. By not allocating objects for non-local sites in the local region, future activations of the method are very likely to have a local region containing only non-escaping objects.

The system is quite straightforward and we have implemented it in Jikes RVM [AAB+00] (see Section 2.1 for the introduction of Jikes RVM). The prototype of the allocator is implemented with the baseline compiler only. When we present the VM-related part in Section 3.3, the stack frame layout refers to the conventions of the baseline compiler.

3.2 Allocator

3.2.1 Heap organization

Various garbage collection techniques have different heap organizations. For example, mark-and-swap collectors use a single space, copying collectors divide space into two semi-spaces, and generational collectors divide the heap into several aging areas. In this chapter, the heap we are discussing refers to the space where new objects are allocated.

A region memory manager organizes a heap as pages. Without loss of generality, the heap in our system is organized as contiguous pages with a fixed size which is a power of 2. The starting address of the heap is aligned to the page size. Therefore, computing the page number for an address requires only subtraction and bit shifting. Some systems do not allocate the large objects from regions, and do not allow an object to straddle two pages. In order to get a full picture of allocation behaviors, our system does not use a separate space for large objects and attempts to allocate objects on contiguous pages whenever it can.

Figure 3.1 gives a high-level overview of the memory organization that we use for our implementation of regions. A page descriptor encodes page status, region identification, allocation point, and the index of next page. A region descriptor contains region status, and the first page index of the region.

This organization provides sufficient information for region-based allocation. When allocating space in a region, the allocator first checks the free bytes of the first page. When there is not enough space left there, a free page is taken from the free list and inserted in the page list as the first page. Allocating space for large objects involves searching for contiguous free pages. We have measured the overhead for these allocations for our benchmarks, and as shown in Section 3.5, the frequency of expensive searches is quite low, indicating that this is a reasonable design.

3.2.2 Services

The internal heap organization is transparent to the JVM. The allocator provides a set of services to the JVM and collector. We describe these functions here.

There are two services for region operations as shown in Figure 3.2. Internally, free region tokens are managed by a stack. The `NEWREGION` service pops a region token from the stack, and pre-allocates one free page for it (pre-allocation is only used with lazy region allocation, to be explained in Section 3.3). If no token is currently available, the GLOBAL one is returned. The `FREEREGION` operation has to check the DIRTY field in the region descriptor. Only when the region is clean, can pages be reclaimed by adding them to the free list. Otherwise, pages are appended to the page list of the GLOBAL region.

```
NEWREGION: int
  if the rid_stack is empty
    return GLOBAL;
  else
    rid = rid_stack.pop;
    pre_allocate_page(rid);
    return rid;

FREEREGION (int rid)
  if the region is dirty
    append pages to the GLOBAL region;
  else
    add pages to the free list;

  rid_stack.push(rid);
```

Figure 3.2: Services for regions

As outlined in Figure 3.3, the allocator provides two services for write barriers. The `CHECKWRITE` service is called before *putfield* and *aastore* byte codes. The addresses *lhs* and *rhs* point to the left hand side and right hand side objects. The operation filters out null pointers and escaped objects first, then computes page indexes from object addresses and tests equality. Region IDs are retrieved from page descriptors and compared if the objects are not in the same page. The *rhs* object is marked as escaped if it is not in the region of the *lhs* object.

The write barrier for *putstatic* calls `MARKESCAPED` directly. As we explained in Section 3.1, the allocator uses a bit vector to record the states of allocation sites. Both services not only mark the region as dirty, but also set the state of the allocation site to *non-local*. In the object header, a bit in the status word is used to mark an object as escaped.

The main function of the allocator is to allocate space for an object. With regions, the allocation of space is somewhat complicated. The allocation process `ALLOC` is illustrated in figure 3.4. Here, we present only a high-level abstraction of the service. The allocation method first checks the state of the allocation site. Only local sites are eligible for allocation from local regions. The internal method *_getHeapSpace* allocates space in the first page if the free space is larger than the required size. If the first attempt fails, it looks at pages following the current page. If the request cannot be satisfied from these pages, it then looks for contiguous pages by scanning the bit maps. This is the most expensive operation in a region-based allocator.

These services also provide the facilities required by the garbage collector to perform collections. We discuss the collection process in Section 3.4.

3.3 Adaptive VM

To utilize regions, a JVM needs the following modifications:

1. each allocation site is assigned a unique index at compilation time;

2. the object header has a field for recording the index of its allocation site;

```
CHECKWRITE (ADDRESS lhs, rhs): boolean
  if rhs is null
    return TRUE;      // case 1

  if rhs is escaped
    return FALSE;     // case 2

  if rhs and lhs are in the same page
    return TRUE;      // case 3

  if rhs and lhs are in the same region
    return TRUE;      // case 4

  mark rhs as escaped,
  return FALSE;       // case 5

MARKESCAPED (ADDRESS rhs): boolean
  if rhs is null
    return TRUE;      // case 1

  if rhs is escaped
    return FALSE;     // case 2

  mark rhs as escaped,
  return FALSE;       // case 3
```

Figure 3.3: Services for barriers

```
ALLOC (int rid, int size): ADDRESS
  call _getHeapSpace(rid, size);

  if failure
    initiate a collection;

    call _getHeapSpace(rid, size);
    if failure
      out of memory;
  else
    return the address;

_getHeapSpace (int rid, int size): ADDRESS
  1. allocate space from the first page;

  2. if failure, check if enough pages
     following the first page are available;

  3. if not, search contiguous pages
     in the free list;

  if both attempts fail
     out of memory;
  else
     add free pages to the region;
     return the starting address;
```

Figure 3.4: Allocating spaces

3. the stack frame has a slot for the region ID at a fixed offset from the frame pointer;

4. the method prologue and epilogue have additional instructions to deal with the region ID slot; and

5. write barriers are inserted before *putstatic*, *putfield*, and *aastore* byte codes.

The allocation method has two more parameters than before: the index of an allocation site is a runtime constant, and the region ID is fetched from the stack frame.

When deciding whether or not a method is eligible for a new local region, our implementation uses following criteria:

- A native call is assigned the GLOBAL region id.

- The <*clinit*> method always uses the GLOBAL region since we know that it initializes static fields.

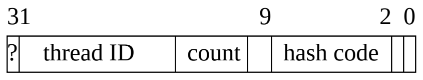
- The <*init*> method inherits the caller's region because it initializes the instance fields.

- A method returning an object is not eligible for a new region. This rule eliminates the need for a write barrier for the *areturn* byte code. More importantly, as pointed out by [GS00], there are many methods just allocating objects for the caller.

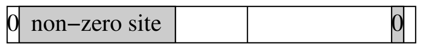
- A one-pass scan of the byte codes counts the allocation sites of each method. If the number is lower than a threshold, no local region is needed for this method. We currently use a threshold of 1.

- The first executed method of the booting thread is assigned the GLOBAL region ID.

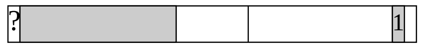
In our initial development it became clear that making *newregion* and *freeregion* calls on each activation is too expensive for the run time system since many activations



(a) Structure of the status word



(b) Status of a non-escaping object



(c) Status of an escaped object

Figure 3.5: Sharing bits with thin locks

may have empty regions, either because their allocation sites have become non-local or because no object is allocated. To eliminate these empty regions, we use lazy region allocation. An eligible method first saves a special region ID, e.g. `0`, in the region ID slot, indicating the stack frame needs a dynamic region, but it has not yet been allocated. The code for allocation first checks the ID, and then calls *newregion* only when necessary. The *freeregion* method is called only when the region ID is a valid one. If a method inherits a region from its caller, it must write back its current region ID to the caller's stack frame.

Another implementation issue is how to encode the allocation site index in the object header. A two-word object header is quite a popular design on most JVMs. One word of the header is used as a status word. Our implementation avoids growing the object header by storing the allocation site index in space already used by the thin lock [BKMS98].

In Jikes RVM version 2.0, the thin lock uses 13 bits for recording the ID of the locking thread, and 6 bits for counting locks. Figure 3.5(a) shows the structure of the status word. Bit 31 is called the *monitor shape bit* which is 0 if the lock is thin and 1 if it is fat.

As indicated in Figures 3.5 (b) and (c), we use bit 1 in the status word to indicate if the object has escaped or not[1]. If the object is non-escaping, then we reuse the thread ID field to store the allocation site (Figure 3.5(b)). This reuse of the thread ID field necessitates some extra machinery for the case where a lock operation is performed on a non-escaping object. The thin lock mechanism first attempts to check the monitor shape bit and the thread ID field in the status word. In ordinary thin locks, the common case is that the monitor shape bit and the thread ID are both zero. However, in our scheme, a non-escaping object is using the thread ID field and it will be non-zero. Thus, when a thin lock fails we must check to see if it failed because a non-escaping object is reusing the thread ID field. If the object is non-escaping, we give back the field to the thin lock by clearing the thread ID field, setting the escaping bit, and then attempting the lock operation again.

By changing a non-escaping object to escaping, we do lose some opportunities for finding local objects, but we do not affect the behaviour of the thin locks. In Section 3.5.5 we show that this effect is not too large. To ensure correctness of this scheme, an escaped object must never become non-escaping, and whenever an object is marked as escaped, the associated region must be marked as dirty.

The system only adds a check on the uncommon path of the thin lock and may need one check on the common path in very few cases. The mechanism allows us to encode allocation site numbers up to $2^{13} - 1$. For large applications, it would be possible to use both the thread ID and lock count to store a 19-bit allocation site index if their positions were reversed (to ensure that small allocation site index still produces a non-zero thread ID field).

There are some other issues related to Java semantics [LY96]. An exception may transfer control to the caller without going through the method epilogue. In this case, the exception mechanism must release the region before unwinding the stack frame.

If an object has a non-trivial *finalize* method, the JVM has to run the finalizer before the space is reused. The region-based allocator organizes the list of objects with non-trivial finalizers by region ID. When the pages of a region are about to be

[1]Bit 1 is used for write barrier purpose in other types of GC. In our prototype implementation in a copying collector, this bit is used as the escaping bit.

reclaimed, the finalize methods of objects in the region get called.

3.4 Collector

The collector must ensure that if an object escapes its original region, that region is marked as dirty. One way of ensuring this would be to introduce write barriers during collections. However, this may sacrifice the efficiency of the collector. There is a trade-off between precision and performance. If all live objects are copied to dirty regions, no barriers are needed. So, the second option is to copy all live objects to the GLOBAL region of another space. This does not violate the above rule since the global region never gets released. This strategy has the same efficiency as a normal copying collector. However, copying all objects to the GLOBAL region may cause some objects created in the next epoch to be treated as escaped, and their associated allocation sites marked as non-local, unnecessarily.

Our current implementation keeps objects in their original region as much as possible, and marks all live regions as dirty after the collection. Now objects in the root set are divided into subsets by their regions, with each live region corresponding to a subset of the root set. The collector starts with collecting all reachable objects from the subset belonging to the GLOBAL region. In the next step, the collector collects objects reachable from the subset corresponding to each local region. All objects copied to the GLOBAL region are marked as escaped to allow fast checks in barriers (the states of the allocation sites are not changed). Although this strategy makes some stackable objects in current live regions unstackable, it does not require write barriers and will not make allocation sites non-local unnecessarily. Currently, we do not have experimental evidence to show which option is better in reality.

3.5 Results

We implemented a page-based heap and a prototype of a region-based allocator in Jikes RVM v2.0 with the baseline compiler. The region-based allocator is implemented in a semi-space copying collector using Cheney's tracing algorithm [JL96]. The implementation uses a uniprocessor configuration. However, it can be implemented in existing parallel collectors with little effort.

To understand the program behavior, we did detailed profiling of the allocation behaviors, and report the experimental results of the following aspects:

- the allocation behavior of the region-based allocator;
- the percentage of space reclaimed by local regions, and the reduction in collections due to local region allocation;
- the behavior of write barriers;
- the impact on thin locks; and
- the effect of adaptivity.

3.5.1 Benchmarks

We report experimental results on the SpecJVM98 benchmark suite and `soot-c`. We first provide some measurements to give some idea of the allocations performed by each benchmark. Table 3.1 shows the profiles of allocation sites. The column labeled *Compiled* gives the total number of allocation sites in compiled methods. It includes the allocation sites in the JVM, libraries and benchmark code. The column labeled *Used* lists the number and percentage of allocation sites which created at least one object. On average 26% of the allocation sites create at least one object. The columns labeled *Non-local* and *Local* show the fraction of used allocation sites which are categorized as *non-local* and *local*. An allocation site is categorized as *local* if it is never marked as *non-local* by the adaptive algorithm. The last column, labeled *Max RID*, gives the maximum number of regions used by the benchmark at the same time. This gives us some idea of the number of region tokens required. Note that a program (like `_213_javac`) with deep recursion may use a large number of regions.

In all of our experiments the total heap size is set to 50M, from which the JVM uses about 1.5M as the boot area. The JVM itself shares the same heap with applications.

Benchmark	Compiled	Used	Non-local	Local	Max RID
_201_compress	2108	346(16%)	115(33%)	231(67%)	11
_202_jess	2407	577(23%)	276(47%)	301(53%)	9
_209_db	2117	358(16%)	124(34%)	234(66%)	11
_213_javac	2871	895(31%)	437(48%)	458(52%)	56
_222_mpegaudio	3266	1502(45%)	157(10%)	1345(90%)	12
_227_mtrt	2228	497(22%)	196(39%)	301(61%)	19
_228_jack	2396	614(25%)	204(33%)	410(67%)	16
soot-c	3030	1158(34%)	551(52%)	507(48%)	14

Table 3.1: Allocation sites

We do not distinguish the objects created by the system or the benchmarks. The heap is divided into two semi-spaces. A 25M heap is quite small for most of our benchmarks, which forces the garbage collector to work.

3.5.2 Choice of page size

The choice of page size may affect utilization of heap space. A larger page size will allow more allocations to be satisfied in the first page. On the other hand, smaller page size will reduce the amount of *froth* (unused pieces of the heap due to the allocation of chunks/pages of memory[2]). Table 3.2 shows the the effect of different page sizes on the number of garbage collections needed and the froth rates. The column labeled *Base collections* gives the number of collections needed for the base semi-space copying garbage collector, without the regions. The three columns labeled *clc* give the collections required for the region-based allocator, assuming page sizes of 256 bytes, 1K bytes and 4K bytes.[3] Similarly, the columns labeled *froth* give the wasted space for the three different page sizes (computed by unused_bytes/allocated_bytes). Note that a page size of 4K leads to a large froth rate for several benchmarks, most

[2]This term was introduced by Steensgaard [Ste00].
[3]We disabled System.gc() calls for both collectors.

notably `_213_javac` (130%), `_228_jack` (27.5%) and `soot-c` (23.5%). The very high froth rate for `_213_javac` also seems to increase the number of garbage collections, which is more than double that of the base collector.

Benchmark	Base #	256 Bytes		1K Bytes		4K Bytes	
	collections	clc	froth	clc	froth	clc	froth
_201_compress	7	7	0.03%	7	0.11%	7	0.47%
_202_jess	12	11	0.13%	11	0.53%	11	2.19%
_209_db	4	4	0.05%	4	0.23%	4	1.05%
_213_javac	12	12	4.96%	15	29.41%	25	130.42%
_222_mpegaudio	0	0	0.62%	0	2.10%	0	9.05%
_227_mtrt	7	1	0.03%	1	0.09%	1	0.38%
_228_jack	9	7	1.29%	8	5.97%	9	27.52%
soot-c	15	13	1.09%	13	4.89%	15	23.49%

Table 3.2: Effect of page size on # of collections and froth

From the perspective of number of collections and froth, the smaller pages seem better. However, this is not the complete story. One must also consider the overhead for allocations. The cheapest form of allocation is when the newly allocated object fits in the current page, the second cheapest is when the allocation can be allocated on the next page, and the most expensive is when one must search the free list for enough contiguous pages to meet the allocation request. These overheads are summarized in Table 3.3. Considering the three page size configurations: 256-byte, 1K, and 4K, the allocations are categorized into three types, which reflect three possibilities in *_getHeapSpace* in Figure 3.3.

1. `firstpage` the space is available in region's first page;
2. `nextpages` the region is expanded with immediately contiguous pages; and
3. `searching` search for contiguous pages in the free list.

A large size (4K) allows most of allocations to be satisfied with cheap costs. But, as we demonstrated in Table 3.2, the froth rate may run out of control. From Table 3.3 we see that there is not a large difference in the behavior of allocations when comparing page sizes of 1K and 4K. However, even though a smaller page size (256 bytes) reduces froth rates, the allocation distribution changes dramatically, with many more allocations requiring expensive operations. From these results we conclude that the trade-off between page size and froth rate is worth considering when using a page-based heap. For our remaining experiments we chose a page size of 1K, which gives us both reasonable froth rate and reasonable allocation overhead.

3.5.3 Region-reclaimed space

The next important measurement is to find out the percentage of space that is reclaimed from local regions. That is, how much space can be reclaimed when using the region-based approach. Recall that the region can be reclaimed when a stack activation is popped only when the dirty bit has not been set (i.e. the region is clean). If any object in the region has escaped, then the dirty bit will be set, and this region must be added to the GLOBAL region which will be collected by the garbage collector.

The table given in Figure 3.6(a) gives the bytes reclaimed from clean regions and the percentage they represent of total allocated bytes, when the page size is 1K. Different page sizes give very similar numbers. The percentage of region-reclaimed bytes varies between benchmarks. In the best case, *_227_mtrt* has 80 percent of total allocated memory reclaimed by regions, with the number of collections reduced from 7 to 1. In the worst case, *_209_db* has less than 1% region-reclaimed space, with no impact on the number of collections.

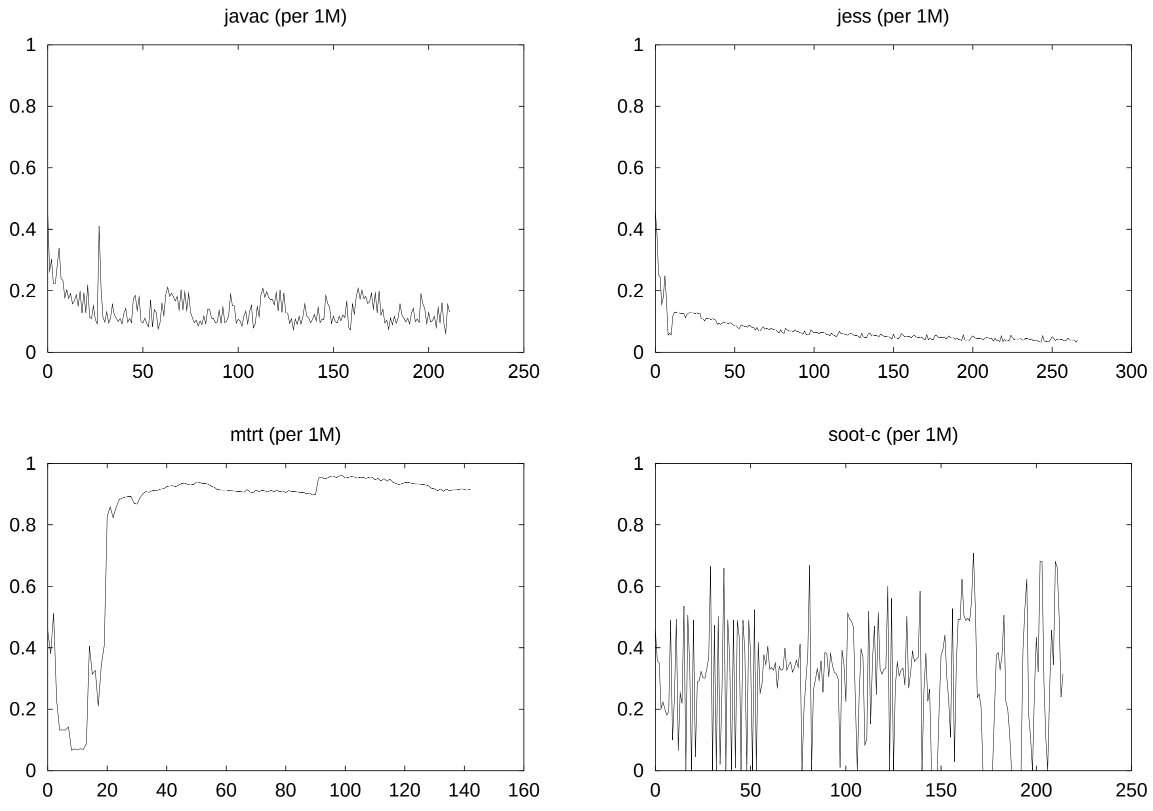
Another way to look at the behavior of the regions is to examine the number of bytes allocated from local regions over the duration of the execution. Figure 3.6(b) shows the fraction of bytes allocated that are allocated from local regions. The x-axis is an abstraction of time, with each unit corresponding to 1M bytes of allocations. The y-axis shows the fraction of those 1M bytes that were allocated from a local region. For example, the graph labeled `_227_mtrt` indicates that after an initial startup, about

Benchmark	page size	firstpage	nextpages	searching
_201_compress	256	82.73%	16.23%	1.04%
	1K	94.96%	4.74%	0.30%
	4K	98.43%	1.44%	0.13%
_202_jess	256	86.75%	12.89%	0.36%
	1K	96.71%	3.27%	0.02%
	4K	99.16%	0.83%	0.01%
_209_db	256	92.24%	7.69%	0.07%
	1K	98.04%	1.92%	0.03%
	4K	99.49%	0.48%	0.03%
_213_javac	256	89.56%	8.58%	1.85%
	1K	97.43%	2.00%	0.57%
	4K	99.41%	0.50%	0.09%
_222_mpegaudio	256	84.97%	12.09%	2.94%
	1K	95.54%	3.52%	0.94%
	4K	98.59%	0.98%	0.43%
_227_mtrt	256	96.04%	2.58%	1.39%
	1K	99.51%	0.38%	0.11%
	4K	99.88%	0.10%	0.02%
_228_jack	256	91.64%	6.46%	1.91%
	1K	97.79%	1.59%	0.63%
	4K	99.48%	0.33%	0.19%
soot-c	256	88.28%	9.85%	1.88%
	1K	96.85%	2.66%	0.49%
	4K	99.21%	0.68%	0.11%

Table 3.3: Allocation behaviors

Benchmark	base collections	total allocated	region-reclaimed	clc	froth
_201_compress	7	116M	15.39M(13.27%)	7	0.11%
_202_jess	12	267M	17.36M(6.50%)	11	0.53%
_209_db	4	77M	0.57M(0.74%)	4	0.23%
_213_javac	12	212M	18.65M(8.80%)	15	29.41%
_222_mpegaudio	0	7M	1.96M(28.00%)	0	2.10%
_227_mtrt	7	143M	115.15M(80.52%)	1	0.09%
_228_jack	9	223M	51.00M(22.87%)	8	5.97%
soot-c	15	219M	40.72M(18.59%)	13	4.89%

(a) Region-reclaimed space



(b) Bytes allocated in local regions

Figure 3.6: Local region behaviors

90% of all allocated bytes are allocated from local regions. In contrast, the graph labeled `_202_jess` shows that this benchmark quickly declines to less than 10% of all allocations from local regions. The graph labeled `soot-c` shows a widely fluctuating rate as the program progresses. This is likely because `soot-c` is quite a complex benchmark with many different phases. It is interesting to note that if all the bytes allocated to local regions are also region-released, then the area under the curves of Figure 3.6(b) should be equal to the region-reclaimed number shown in Figure 3.6(a). This appears to be the case, as confirmed by our measurements given in Section 3.5.6, showing that almost all local regions are clean when released.

3.5.4 Write barrier behaviors

Another important aspect of the collector we measured is the behavior of write barriers, as summarized in Table 3.4. Overall, a write barrier has bounded constant time as shown by the pseudo-code in Figure 3.3. However, it is still a burden to the system. To get better idea of how to optimize the barriers, we categorize the `CHECKWRITE`(CW) for *putfield* and *aastore* into five types.

1. the right hand side is a null pointer;
2. the right hand side object is already marked as escaped;
3. both objects are in the same page;
4. both objects are in the same region; and
5. the *LHS* and *RHS* objects are not in the same region.

The first case checks if the right hand side reference is a null pointer, and the second case checks the escaping bit, which requires a load and a compare instruction. As shown in Table 3.4, the majority of checks are filtered out by these two cases, which indicates it is beneficial to separate these two cases as a common path and inline them. The remaining three cases can be processed by a method call. Similarly the write barriers for *putstatic* are also categorized into three types, with the first two cases benefiting from inlining.

Benchmark		null	quick	samepage	sameregion	escaped[1]
_201_compress	CW[2]	15.07%	83.64%	1.05%	0.17%	0.07%
	MS[3]	0.00%	91.45%			8.55%
_202_jess	CW	0.22%	99.75%	0.02%	0.00%	0.00%
	MS	1.88%	85.63%			12.50%
_209_db	CW	0.12%	99.88%	0.00%	0.00%	0.00%
	MS	0.00%	90.48%			9.52%
_213_javac	CW	10.81%	88.84%	0.32%	0.03%	0.00%
	MS	2.90%	89.63%			7.47%
_222_mpegaudio	CW	0.42%	98.12%	0.08%	1.38%	0.00%
	MS	3.20%	83.56%			13.24%
_227_mtrt	CW	13.82%	84.49%	1.66%	0.02%	0.00%
	MS	0.00%	86.44%			13.56%
_228_jack	CW	11.79%	87.58%	0.62%	0.01%	0.00%
	MS	0.00%	91.00%			9.00%
soot-c	CW	5.20%	91.88%	2.71%	0.21%	0.00%
	MS	0.00%	93.48%			6.52%

[1]In this table, zero only means the rate is lower than 0.005%. [2]CW is the short name of CHECKWRITE for `putfield` and `aastore`. [3]MS is the short name of MARKESCAPED for `putstatic`.

Table 3.4: Write barrier behaviorss

Currently, we have not incorporated any static analysis for removing write barriers. Certainly, such analyses will reduce the runtime cost of the system. We will investigate such algorithms in the future.

3.5.5 Impact on thin locks

We also profiled the impact of sharing bits with thin locks. Table 3.5 shows the rate of failed locks because of sharing bits. *_201_compress* and *_222_mpegaudio* have only a few thousand locks in a full run, so their results cannot represent the real effect of sharing bits. On other benchmarks, the rate of spoiled locks is no more than 5%.

Benchmark	thin locks	spoiled locks
_201_compress	1.6K	172(9.58%)
_202_jess	4.8M	4881(0.10%)
_209_db	45.2M	2915(0.01%)
_213_javac	14.7M	341585(2.26%)
_222_mpegaudio	5.6K	497(8.15%)
_227_mtrt	1.3M	25816(1.92%)
_228_jack	9.4M	497016(5.00%)
soot-c	5.6M	73204(1.30%)

Table 3.5: Impact on thin locks

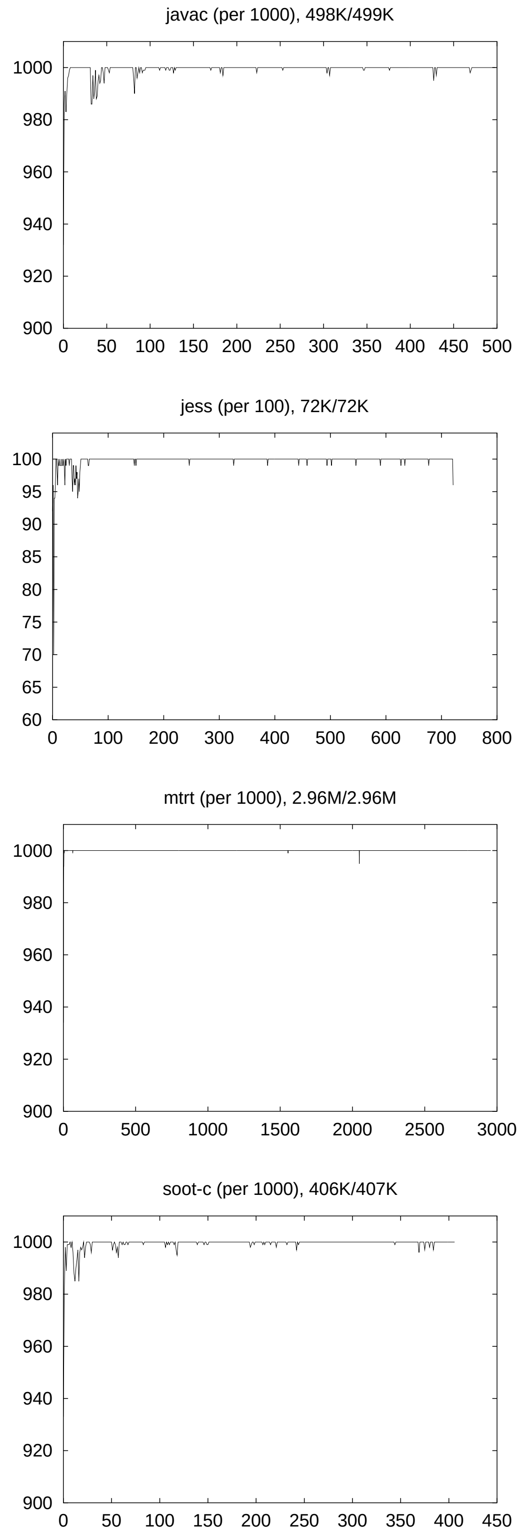
3.5.6 Effectiveness of adapting

The last behavior that we studied was the effect of the adaptive part of our algorithm. The basic idea of our approach was to mark allocation sites as non-local as soon as they are found escaping the first time. The justification of this decision was that this would prevent this allocation site from spoiling clean regions in the future, and we expected that this would lead to most local regions being clean at release time. Figure 3.7(a) shows the number of local regions that are clean at release time over the duration of the execution of the program. The x-axis is an abstraction of time, with each point representing the release of 1000 local regions (100 for `_202_jess`). The y-axis shows the number of those local regions that are clean at release time. Accompanying the title of each graph is the number of clean regions and total allocated regions. It is very clear that after a short startup time, the system quickly adapts so that almost 100% of the local regions are clean at release time. There is occasionally a small dip, but then the system adjusts and it goes back to almost 100%. So, it does appear that the system adapts well.

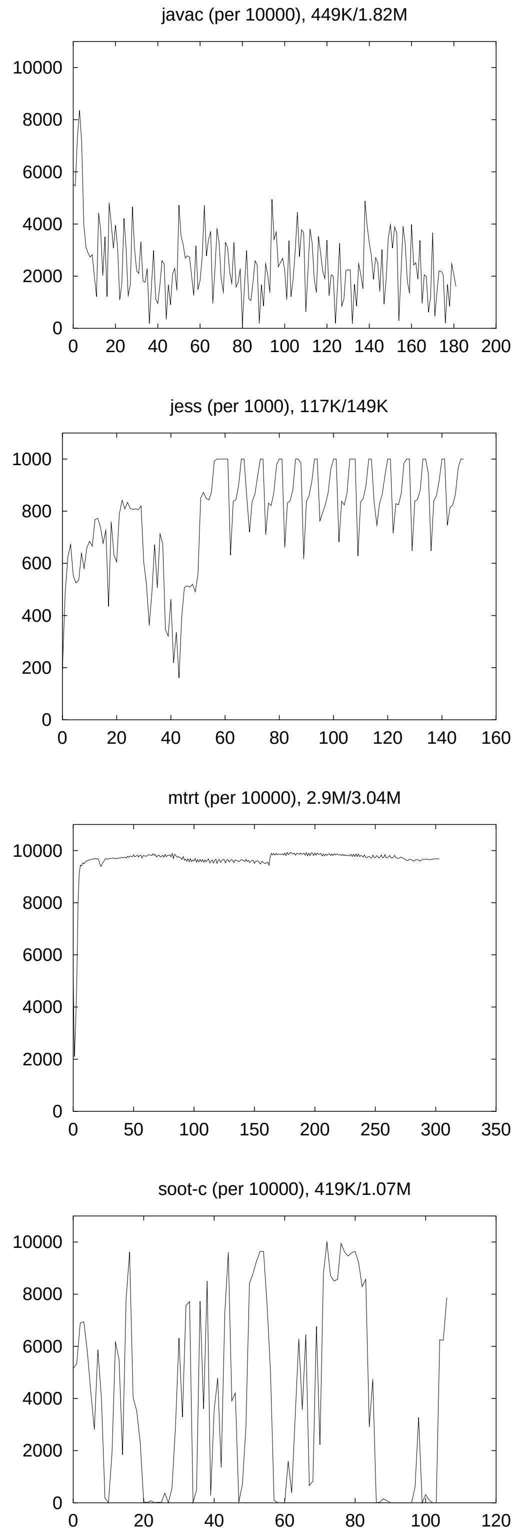
In order to see what would happen without adapting, we removed the part of the algorithm that marks an allocation site as non-local, so that **all**[4] allocations are placed in the local regions. Figure 3.7(b) shows the result in this case. First, note that many more regions are created, and the scale on these graphs are now per 10000 local regions (1000 for `_202_jess`). However, we can also see some interesting trends. The benchmark `_227_mtrt` appears to create mostly non-escaping objects, so for this benchmark it is not such a bad idea to just put all objects in local regions. For benchmarks `_213_javac` and `soot-c`, we see that removing the adaption leads to many more regions, and many of those regions are not clean. In these cases the adaption works to cull those dirty regions. For `_202_jess` we see a very interesting behavior, in that in the last two thirds of the execution, a lot of objects appear to be non-escaping, and the number of clean regions stays quite high. With the adaption, we get a higher percentage of clean regions, but we don't find nearly as many. In this case we suspect that there are some allocation sites which sometimes produce escaping objects, and sometimes not. A more complicated prediction scheme appears to be necessary for this kind of benchmark.

We also collected our overall measurements for the two cases, with adaption and without adaption. These are summarized in Table 3.6. Note that in some cases, most notably `_213_javac`, switching off the adaption drastically increases the froth rate (589% instead of 29%) and number of garbage collections (96 instead of 15). However, as we might have predicted from the graph in Figure 3.7(b), the performance

[4]except native calls, <*clinit*>, and the first executed method of the boot thread, see Section 3.3.



(a) With adaption



(b) No adaption

Figure 3.7: Clean regions released

for _202_jess is much better without adaption. This is a clear sign that we must look at other forms of adaption that are more robust when the objects created from a site sometimes escape, and sometimes do not. Overall, it seems that the adaptive algorithm gives better performance and controls excessive froth.

The third column of Table 3.6 gives the total size of escaping objects which were captured by write barriers or reachable from escaping objects. Locked objects are also considered as escaping. The difference between the total allocated size and the size of escaping objects gives us a rough upper bound of the space that can be reclaimed by regions. We see that there is a large space to improve the current prediction scheme.

Benchmark	total allocated	total escaped	base clc		region reclaimed	clc	froth
_201_compress	116M	99M	7	WA[1]	15.39M (13.27%)	7	0.11%
				NA[2]	14.75M (12.72%)	7	3.61%
_202_jess	267M	6M	12	WA	17.36M (6.50%)	11	0.53%
				NA	224.92M (84.24%)	2	9.04%
_209_db	77M	24M	4	WA	0.57M (0.74%)	4	0.23%
				NA	0.53M (0.68%)	4	5.77%
_213_javac	212M	112M	12	WA	18.65M (8.80%)	15	29.41%
				NA	24.41M (11.51%)	96	589.09%
_222_mpegaudio	7M	2M	0	WA	1.96M (28.00%)	0	2.10%
				NA	1.37M (19.57%)	0	128.43%
_227_mtrt	143M	17M	7	WA	115.15M (80.52%)	1	0.09%
				NA	112.56M (78.71%)	6	62.42%
_228_jack	223M	69M	9	WA	51.00M (22.87%)	8	5.97%
				NA	94.66M (42.45%)	7	16.26%
soot-c	219M	89M	15	WA	40.72M (18.59%)	13	4.89%
				NA	7.82M (3.57%)	57	276.54%

[1] WA is the short name for With Adaption. [2] NA is the short name for No Adaption.

Table 3.6: Effect of Adaption

3.5.7 Summary

Our current implementation, using the baseline compiler, was aimed at producing a prototype that could be used to measure the behavior of the system, as we have presented in this section. Our numbers show that: 1) page size is important, but with the appropriate page size, the overhead for froth and the frequency of expensive searches for free pages is quite low; 2) for many benchmarks a significant percentage of allocated memory can be placed in local regions which are still clean at release time; 3) appropriate choices for the barrier operations can put the common cases on a low-cost path; 4) the overhead for sharing space with thin locks seems acceptable; 5) the adaptive part of the algorithm is important for focusing the system on the allocation sites that are likely not to escape; and 6) for many, but not all, benchmarks the adaptive system finds more local regions than a non-adaptive system.

We did try measuring runtime improvement using this prototype, but it turned out that the overheads in our current implementation are still too high, and this can lead to an overall slow down. This is due to several factors, including: 1) the cost of region management, 2) the cost of write barriers, and 3) small helper methods used by region and barrier implementation.

In our preliminary experiment with the optimizing compiler, we found that the optimizing compiler allocates far more objects than benchmarks. It is hard to measure the effects of regions on applications.

3.6 Related work

We have described a region-based allocator using page-based heaps. Although we use the terminology *region* here, the technique does not involve any region inference algorithm [TT97]. The technique provides an alternative way to allocate objects on stack in a JVM. There is much literature on garbage collection, region-based memory management and object stack-allocation, thus we focus on those systems most closely related to our work.

Tofte's region inference system [TT97, Tof98] automatically infers regions for objects. It achieves automatic memory management by compiler analyses. Gay and Aiken's C@ [GA98] and RC [GA01] provide language support for regions. C@ does not require an inference algorithm. It uses reference counting and stack scans to determine the safety of reclaiming a region. The main point of our work was to develop a system that works for an existing language, Java, and that is transparent to the Java programmer. Steensgaard [Ste00] proposed thread-specific heaps for multi-threaded programs. Both systems require the heap to be organized as pages/chunks. We studied the allocation behaviors of Java programs on page-based heaps. The preliminary results suggest that Java programs are sensitive to the page size.

Escape analyses [CGS+99, Bla99, WR99, GS00] for Java determine whether the objects created by an allocation site may escape certain scopes. Mainly the analysis results can be used in two optimizations. Thread escape analysis results can be used to remove unnecessary synchronizations, and escape analysis to find method-bounded allocation sites can be used to create objects on the local stack frames. However, the cost of the analyses prevents them from being used at run time, and Java semantics may pose restrictions on stackable objects. Our region-based allocator aims to reduce the work of garbage collector by allocating objects in temporary regions. The technique needs no analyses and may be suitable for a run time system like a Java Virtual Machine. One of our future research directions is to explore online escape analysis to remove unnecessary barriers for region allocation.

McDowell [McD] reported the number of potentially stackable objects in a set of Java benchmarks. Like other escape analyses, McDowell also made the assumption that a compile time algorithm must make a decision for all objects created by an allocation site, although he was using dynamic profiling information to conclude the results. Our system does not require this limitation. An allocation site may create objects in local regions before any of them is found to be escaped. Extensions of our adaption algorithm may also allow allocation sites to become local again, even after being marked as non-local.

Hallenberg [Hal99] introduced garbage collection into individual regions in Tofte's region inference system for the ML Kit. Although our collector has a similar name as his system, the structures are quite different. In his system, the region inference algorithm creates regions, and inside a region, a copying collector collects live objects. The backbone of our system is a garbage collector, and the region is a natural way to extend stack frames. The region organization serves as the basis of adaptive allocation. Interested readers can find the design of Hallenberg's system in [Hal99], chapter 11.

3.7 Conclusions

We have presented an adaptive, region-based allocator for Java Virtual Machines and studied the allocation behavior of Java programs on page-based heaps. The main idea is to detect on-the-fly these allocation sites that do not escape their region, and then manage these allocations in local regions that can be released when the associated stack frame is popped.

We implemented the system using the Jikes RVM baseline compiler and associated garbage collector, and we used this prototype to study the behavior of a collection of Java benchmarks, including the SpecJVM98 benchmarks. This study showed that the design of the system is crucial, including an appropriate choice of page size, and techniques for minimizing space overhead and region allocation/deallocation overhead. We also studied the adaptive mechanism of our system, and found that it quickly found regions from which no object escaped.

Given our encouraging results, there are several directions to continue on this work. First, we would like to profile and analyze important escaping allocation sites, and develop necessary techniques to allow more objects be allocated on regions.

Our second major area of investigation is to look at a wider range of adaptive mechanisms. In this work, we mark an allocation site as non-local as soon as one object allocated from that site escapes. This scheme is a rather coarse and naive prediction scheme. When we turn off the adaptation, this corresponds to a predictor that always predicts that no object will escape. Our experiments show that this second method works well for those programs where, in fact, a large portion of the objects do not escape. We would now like to examine other more complex predictors. For example, we could use more than 1 bit, and only mark an allocation site as non-local after some number of objects escape. We could also increase the granularity of the predictor by associating dirty bits with pages within regions, rather than having 1 bit per region. Another possibility is to reset allocation sites to being local at intervals, for example at collection time or during phase shifts in the program. This may help in the case where the same allocation site is sometimes local and sometimes non-local. As part of this study we also would like to measure the effect of the regions on memory locality.

Our final area of research is to examine the effect of our system when coupled with different garbage collectors. As we pointed out in Section 3.4, it should be relatively straightforward to incorporate our ideas in a variety of collectors.

Chapter 4
Runtime Techniques Supporting Speculative Optimizations

Static program analyses for optimizations must be conservative to preserve the semantics of original programs. Data-flow facts induced by static analyses must be held regardless input data or execution environments because once the program is compiled to binaries, the compiler has no control on the code and execution. In a JIT environment such as a Java virtual machine, since compilers are part of the execution engine, it is possible to dynamically change the compiled code.

In a Java virtual machine, classes are loaded dynamically, and methods are compiled lazily. When compiling a method, program properties collected by compilers are held for loaded classes and compiled methods, but might be invalidated when more classes are loaded later. We call optimizations based only on compile-time properties *speculative*. Speculatively optimized code is only safe at compile time and might be wrong in future execution.

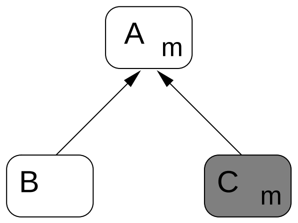
In this chapter, we review existing techniques for supporting speculative optimizations in Java virtual machines. First, we review runtime guards for method inlining in Section 4.2. Section 4.3 introduces more complicated code rewriting techniques: code patching and on-stack replacement. Finally, in Section 4.4, we present our improvement and implementation of an on-stack replacement in Jikes RVM. These techniques can be used for supporting speculative optimizations using dynamic interprocedural analyses results. Section 4.5 discusses related work.

4.1 Speculative inlining

Methods are important semantic abstraction boundaries of object-oriented programs. Compiler writers, however, work hard to remove these boundaries to improve performance. In the Java programming language, a virtual call has the form of <`x, A.m()`> where `x` is a variable pointing to objects of type `A` or its subtypes, and `m()` is the method signature. The real target of each invocation depends on the type of object pointed to by the variable `x` at runtime. We call the object pointed to by `x` the *receiver* of the call, and `m()` is a *message* sent from the caller. A method call requires setting up a stack frame, passing parameters and returning the result. A virtual call requires extra costs of looking up the target method. Method inlining is an important optimization of object-oriented programs. Inlining reduces the direct cost of calls and creates new optimization opportunities for the inlined code.

Due to polymorphism, a virtual call may invoke several different methods in the course of program execution. The target method is determined by the type of the receiver object and the callee method signature. A virtual call site can be categorized into *polymorphic* or *monomorphic*. A polymorphic call site has more than one target method during execution, and a monomorphic call site has only one target even it is virtual. A type analysis, such as class hierarchy analysis (CHA) [DGC95], can prove some virtual call sites are monomorphic and a compiler can inline these monomorphic call sites *speculatively*. However, dynamic class loading grows the class hierarchy during the program execution, and it may invalidate previous CHA results for inlining. For example, a virtual call formally recognized as monomorphic by CHA might become polymorphic in the future. Speculative inlining requires a backup mechanism to ensure the correctness of optimized code in the presence of dynamic class loading.

We use an example in Figure 4.1 to illustrate speculative inlining. A class `A` declares a virtual method `m()` and a class `B` extends `A` without overriding the method `m()`. Another class `C` extends `A` but overrides the method `m()`. We use the class name followed by the method name (e.g., `A.m()` and `C.m()`) to distinguish methods in different classes.



```
if (cond)
  foo(new B());
else
  foo(new C());

void foo(A a) {
  ......
  a.m();
  ......
}
```

Figure 4.1: An inlining example

If class `C` is not loaded when the method `foo` gets recompiled, CHA concludes that the receiver type set of the call site `a.m()` is only `A` and `B` (assuming both were loaded). The compiler can resolve the call target to be `A.m()` only. Thus, this is a monomorphic call at the moment of optimization, and the compiler inlines `A.m()` into `foo`. However, the compiler needs a mechanism to ensure that only objects of type `A` or `B` can reach inlined code of `A.m()` since the class `C` can be loaded in the future execution.

4.2 Class and method tests

4.2.1 Guarded inlining using class and method tests

First we review a simple technique, *class tests*, for ensuring the correctness of speculative inlining. In Jikes RVM, as shown in Figure 2.1, an object has a pointer to its class's *type information block* (TIB). The compiler generates the code for the virtual call `a.m()` of the form:

```
a_tib = a->TIB;
if (a_tib == A's TIB)
  inlined A.m()
else
  normal virtual call
```

The inlined code of `A.m()` is guarded by a *class test*. If the runtime type of `a` is `A`, the control falls through into inlined code of `A.m()`. Otherwise, it goes to the more expensive virtual call implementation.

The drawback of *class test* is that it only covers the case when an object type is `A`. If the object type is `B`, the control falls to the normal virtual call, although the target is still `A.m()`. One remedy is to change the test to `a_tib == A's TIB || a_tib == B's TIB`, but it has two tests for type `B` and increases code density.

Detlefs and Agesen [DA99] provided a new solution to the dilemma of class test. It uses *method tests* to guard inlining of virtual method invocations with the assumption that the target can be obtained from the class information cheaply. In most Java virtual machines, obtaining a method address from a virtual method table takes only one load instruction. *Foo*'s compiled code using method tests would be as follows:

```
a_tib = a->TIB;
m_addr = a_tib[m->method_offset];
if (m_addr == A.m's address)
  inlined A.m()
else
  normal virtual call
```

A single method test can cover more classes than a class test with the cost of one load instruction in the fast path. The two techniques can be selectively used in practice. For example, if the inlined method is in a `final` class, the class test can be used since there is no other receiver types due to language constraints.

4.2.2 Direct inlining using preexistance analysis

Class and method tests are mechanisms guarding inlined code safely even in the presence of dynamic class loading. However, both tests require memory accesses (to load TIBs and method instruction addresses) and conditional branch instructions. For monomorphic calls, these tests seem redundant except as assurance. Detlefs and Agesen [DA99] pointed out that, in a Java virtual machine, a compiler can remove the tests for currently monomorphic calls, and register the compiled code with the assumption that the inlined method is not overridden (the call is monomorphic). Right after a class loader loads a new class, it checks all methods of superclasses overridden by methods in the newly loaded class. For compiled methods whose assumptions were invalidated, their entries in virtual method tables are conveniently reset to lazy compilation code stubs (see Chapter 2.1). The next invocation of an invalidated method triggers recompilation using the new, correct CHA results. In our example, loading class `C` causes resetting `foo`'s machine code address, and the next invocation of `foo` triggers the recompilation of `foo` without direct inlining of `a.m()`.

However, there is a drawback in this approach. Only the next invocation of an invalidated method can trigger recompilation and execute on the correct code. If a thread is executing a compiled method while dynamic class loading invalidates it, resetting the virtual method table entry cannot correct the execution of the thread. Consider the code:

```
foo_1(A a) {
  ......
  if (cond) a = getC();
  a.m();
  ......
```

If the condition has never been satisfied before and `C` is not loaded, `a.m()` is directly inlined into `foo_1` using `A.m()`. If, later on, the condition is satisfied, execution of `getC()` triggers loading of `C` which invalidates the inlining assumption of `A.m()`. Resetting the virtual method table entry of `foo_1` only corrects future invocations, but the current execution of `foo_1` is apparently wrong.

Detlefs and Agesen [DA99] pointed out that, if the receiver object of a call is *preexisting* before the method, resetting its entry in virtual method table is sufficient to ensure the safety because these method activations running on threads would not create objects of newly loading classes. They proposed a simple *invariant argument analysis* for proving the preexistance of receivers. The analysis tracks simple data-flow between reference type variables. If a variable is only assigned values from parameters, then it is preexisting before the method call. In our example, the parameter `a` of `foo` is preexisting, but `a` of `foo_1` is not.

For monomorphic calls whose receivers could not be proved to be preexisting, *thin guards* [AR02] can be used to combine several tests into one test. Although guarded inlining removes the overhead of building call stacks and passing parameters, the control flow created by guards limits the effectiveness of further optimizations on inlined code. The merge of inlined code and backup path essentially removes the data-flow benefits from inlining. Thin guards are able to create large regions of inlined code for optimizations. Control-flow splitting [CU91] breaks the merge by duplicating the path and improves the effectiveness of optimizations.

4.3 Code patching and on-stack replacement

In this section, we discuss more complicated techniques for direct inlining of monomorphic virtual calls in the presence of dynamic class loading. One technique, code patching, rewrites speculatively inlined code if dynamic class loading invalidates the inlining assumption. Another technique, on-stack replacement, dynamically changes the activation of an invalidated method to a safe one.

4.3.1 Code patching

Preexistance analysis may not always succeed on removing method and class tests for inlined monomorphic calls. Ishizaki et. al. presented a code patching technique [IKY$^+$00] to remove all method and class tests for monomorphic calls in the presence of dynamic class loading. Code patching uses dynamic CHA to identify monomorphic calls based on the class hierarchy at compile time.

Like a method or class test, code patching uses a guard at compile time. However, the guard is not a test, but a condition that the inlined method is not overridden. Using the previous example, Figure 4.2(a) shows the pseudo-IRs of inlined call. If `A.m()` is not overridden at compile time, the compiler directly inlines it into `foo`, and generates a backup path for the call. The conditional test is treated as a normal conditional branch instruction until code generation phase. When generating machine code for the test, it is replaced by a label, *start_of_inlined_code*, as shown in Figure 4.2(b), and the condition, *A.m() is overridden*, is registered in a database together with the machine code address offsets at *start_of_inlined_code* and *backup*. If `C` is loaded later and `C.m()` overrides `A.m()`, the Java virtual machine immediately patches the first instruction at the label *start_of_inlined_code* by a direct jump to the *backup* path as shown in Figure 4.2(c). The virtual machine can optionally reset `foo`'s entry in the virtual method table.

The advantage of code patching is that it removes memory loads and conditional branch instructions. Also recompilation is not a necessity. The disadvantage is that it needs to track detailed dependency and synchronize caches explicitly on some architectures. Preexistance based inlining can still be used together with code patching.

4.3.2 On-stack replacement

In a Java virtual machine, it is convenient to transition to a newly compiled version of a method by resetting the entry of virtual method table. So the future method invocations branch to the new version. However, the transition for a method that is currently executing on some thread's stack presents a harder engineering challenge. In the case of direct inlining, if an invalidated method is executing on a thread's stack, the method activation must be transferred to a version without direct inlining. Preexistance based inlining avoids direct inlining of calls that might be invalidated and lead to wrong execution. Code-patching technique uses a similar idea as guarded inlining except it removes memory accesses and conditional branch instructions. Now

```
if (A.m() is overridden) goto backup;
  inlined A.m()
  ......
backup:
  normal virtual call of a.m()
```

(a) compile-time guard of code patching

```
start_of_inlined_code:
  inlined A.m()
  ......
backup:
  normal virtual call of a.m()
```

(b) inlined call

```
start_of_inlined_code:
  jump backup
  ......
backup:
  normal virtual call of a.m()
```

(c) after code patching

Figure 4.2: Code patching example

we look at a more general and advanced technique, on-stack replacement, for invalidating speculatively optimized code without the preexistance requirement, while retaining better data-flow effects of optimized code than guarded inlining.

To perform the transition to a new version of compiled code, the SELF programming language implementations [Höl94] pioneered on-stack replacement (OSR) technology. OSR technology enables fundamental VM enhancements, including debugging optimized code via de-optimization [Urs92], deferred compilation to improve compiler speed and/or code quality [CU91], online optimization of activations containing long-running loops, and optimization and code generation based on speculative program invariants.

Figure 4.3(a) is an example for demonstrating the concept of on-stack replacement. In this case, the variable `a` is not preexisting, but `A.m()` is currently not overridden. As shown in Figure 4.3(b), the compiler can perform direct inlining by inserting an *osr_point* in the code before the inlined code and after the assignment of `a`. An OSR point is an intermediate instruction that uses all live variables at the point in the original program. The OSR point also keeps a map from variable names to values or locations after register allocation. In the example, the *osr_point* instruction records the stack locations of variables `x` and `a`, and also records the constant values of variables `i`, `j`, and `k`. If the class `C` is loaded later on, the class loader can change the instruction at *osr_point* to a call of a special function to perform on-stack replacement. The transition of optimized (wrong) code to unoptimized (correct) code is shown in Figure 4.3(c). The left side is the state of `foo`'s optimized code when the execution reaches the *osr_point*. The right side is reconstructed stack frame for unoptimized code at the *entry_point*. The machine program counter is set to the *entry_point* and the execution will continue on unoptimized code with the new stack frame.

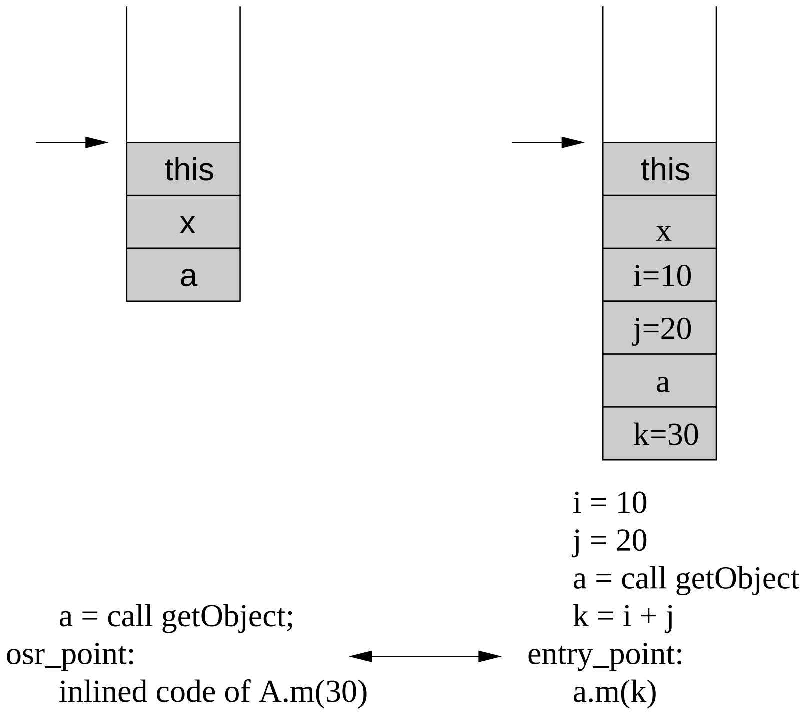
Hölzle et. al. proposed an approach [Urs92] for performing transition for SELF programs. Their approach introduces *interrupt points* in code to be invalidated and *entry points* in the middle of another version of code. To perform transition from one version of code to another version, the first step is to recover the source-level program state. It is straightforward to recover such states from unoptimized code. For optimized code, the compiler inserts a few *interrupt points* which are equivalent

```
void foo(int x) {
  int i = 10, j = 20;
  A a = getObject();
  int k = i + j;
  a.m(k);
  ......
}
```

(a) source code

```
a = call getObject;
osr_point (x, i=10, j=20, k=30, a);
inlined code of A.m(30);
......
```

(b) direct inline of A.m();



(c) transition to unoptimized code at OSR point

Figure 4.3: On-stack replacement example

to OSR points shown in above example. Each interrupt point has a *scope descriptor* which is used to reconstruct the source-level state. The second step is to build a stack frame for the second version of the code at the entry point where interrupt point happens. It is also straightforward to build the stack frame and make transition to unoptimized code. For optimized code, the compiler pre-determines entry points (which are usually at the back edge of the control flow), where the stack layouts and machine code addresses are recorded. A *transition function* does state recovering, new frame construction, and the program counter adjustment. Indeed, Figure 4.3(c) is a very close example of SELF's interrupt point.

Sun's HotSpot server compiler [PVC01] implements SELF's approach for deoptimization (from optimized code to unoptimized code) and promotion of long-running loops (from unoptimized code to optimized code).

Fink and Wegman [FW00] invented a new transition mechanism for performing on-stack replacement. The key idea is that, instead of pre-inserting *entry points* into the target version of code, one can generate specialized source code (bytecode) to set up the new stack frame(s) and continue execution at the desired program counter. The mechanism can be best illustrated by the example shown in Figure 4.4.

Figure 4.4(a) is the Java source code of method `foo` and the bytecode is shown in Figure 4.4(b). The execution state of a method is represented by a *JVM scope descriptor*, as shown in Figure 4.4(c), which is comprised of:

1. the thread running the activation,
2. the program counter as a bytecode index,
3. values of live local variables and stack locations, and
4. a reference to the activation's stack frame.

Given the *JVM scope descriptor* of a method, Fink's approach constructs a special method, in bytecode, that sets up the new stack frame and continues execution (see Figure 4.4(d)). The special method is only used once to complete the execution of replaced method activation, and it preserves the correct semantics of the program.

```
void foo(int x) {
  int i=10, j=20;
  A a = getObject();
  int k = i + j;
  a.m(k);
  ......
}
```

(a) Java source code

```
running thread: MainThread
frame pointer: 0xSomeAddress
program counter: 21
local variables: L0(this), L1(x), L2(i)=10,
                 L3(j)=20, L4(a), L5(k)=30
stack expressions: S0=a, S1=30
```

(c) JVM scope descriptor for an activation of `foo`

```
 0: bipush 10
 2: istore_2
 3: bipush 20
 5: istore_3
 6: aload_0
 7: invokevirtual getObject()LA;
10: astore 4
12: iload_2
13: iload_3
14: iadd
15: istore 5
17: aload 4
19: iload 5
21: invokevirtual A.m(I)V
......
```

(b) bytecode

```
    load (this)
    astore 0
    ldc (x)
    istore 1
    ldc 10
    istore 2
    ldc 20
    istore 3
    load (a)
    astore 4
    ldc 30
    istore 5
    load (a)
    ldc 30
    goto 21

 0: bipush 10
    ......
21: invokevirtual A.m(I)V
    ......
```

(d) specialized version

Figure 4.4: Example of Fink's OSR approach

Fink's approach does not need a transition function to set up target stack frame and does not require *entry points* in the target code. However, it needs to construct and compile a specialized method. SELF's approach requires modifications in compilers to generate *entry points* in the code, which may limit the effectiveness of some optimizations and require large engineering efforts. In contrast, Fink's approach only requires small modifications on existing compilers to compile specialized methods. The specialized method initializes locals and stacks with runtime constants, this could enable more optimization opportunities. The cost of the Fink's approach is that it requires creating one specialized method for one replacement, and increases the workload of dynamic compilation system.

We have extended Fink and Wegman's mechanism to support de-inlining, and implemented it in Jikes RVM (in collaboration with one of original authors, S.J.Fink). Based on this OSR implementation, several optimizations, such as long-running method promotion and deferred compilation, have been developed and evaluated [FQ03].

4.4 Improvement and implementation of on-stack replacement in Jikes RVM

When the optimizing compiler compiles a method, it first determines where to insert *osr_points* in the IR. The decision solely depends on the application of OSR. An OSR point indicates the bytecode-level state can be recovered and on-stack replacement can happen when the program execution reaches the program point.

4.4.1 Transition to de-inlined code

We extended Fink's approach to support *de-inlining*. The ordinary approach can handle the OSR points in the caller's code. When the compiler inlines a call, an *osr_barrier* instruction is inserted before the inlined call site. The barrier collects live variables before the call instruction and is passed to the compilation context of the inlinee. If an *osr_point* instruction is inserted in the inlined code, it uses not only the inlinee's live variables, but also those in *osr_barrier* instructions from the caller. This provides enough information to perform *de-inlining* if the *osr_point* in the inlined code is reached. After register allocation, the compiler builds a table mapping the lives to physical positions (register numbers or spilling offsets) with detailed type information. The table is encoded as an OSR map, some auxiliary information of the compiled method.

When on-stack replacement happens, the system recovers *JVM scope descriptors* for methods in the inlining chain. Each descriptor except the leaf refers to its inlinee's descriptor. Figure 4.5(a) shows an example where method `bar` was inlined into method `foo`, and on-stack replacement happens at the label `B` of `bar`. Recovered JVM scope descriptors for `foo` and `bar` are shown in Figure 4.5(c) where `foo`'s descriptor has a reference to the inlinee `bar`'s descriptor. As shown in Figure 4.5(c), the specialized bytecode for the inliner, `foo_prime`, in addition to set up local variables and stack expressions, makes a call to the specialized bytecode of the inlinee, `bar_prime`. Right after the call, the control jumps to the next instruction of the original call. In `bar_prime`, the prologue sets up locals and stacks.

4.4.2 Implementation

We have fully implemented Fink's approach and improvement in Jikes RVM with applications of deferred compilation, long-running loop promotion [FQ03]. Linux/x86, Linux/PPC, and AIX/PPC are supported. Here we discuss the general implementation strategy and engineering challenges encountered and our solutions.

When on-stack replacement is used for promoting long-running loops, the compiler needs to extract the *JVM scope descriptor* from unoptimized code, which is relatively easy because the baseline compiler generated code that mimics the JVM stack machine. If an application requires deoptimization, it needs help from the optimizing compiler to insert *osr_points* in the code. In the optimizing compiler, an *osr_point* is implemented as an *OsrPoint* instruction in Jikes RVM's IR, which uses live variables at its insertion point. An *OsrBarrier* is an instruction to hold live variables before an inlined call. Live variables are aggregated to *OsrPoint* instructions in inlinees. After

```
void foo() {
  bar();
 A:
  ......
}

void bar() {
  ......
 B:
  ......
}
```

(a) Java code. *Bar* is inlined into *foo* and an OSR happens at program point `B`.

```
foo's descriptor:
  running thread
  ......
  frame pointer
  reference to bar's descriptor

bar's descriptor:
  running thread
  ......
```

(b) JVM scope descriptors for *foo* and *bar*

```
foo_prime :
  <specialized foo prologue>
  call bar_prime
  goto A;
  ......
  bar();
 A:
  ......

bar_prime :
  <specialized bar prologue>
  goto B;
  ......
 B:
  ......
```

(c) special methods for OSR transition

Figure 4.5: De-inlining example

parsing inlined code, *OsrBarriers* are removed. As we discussed before, places to insert OsrPoint instructions solely depend on the OSR application. For example, in deferred compilation, *OsrPoints* replaces branch targets or exception handlers which are never executed. An *OsrPoint* is also a GC point for reusing the code to build GC maps.

Before generating machine code, *OsrPoint* instructions are expanded to runtime services. Currently there are two implementations. The first implementation uses the adaptive compilation system in Jikes RVM. An *OsrPoint* is implemented as a special thread switch point. When a thread hits an *OsrPoint*, it is suspended. A separate compilation thread inspects the suspended thread, performs on-stack replacement for it, and re-schedules it afterwards. The second implementation simulates lazy method compilation: each *OsrPoint* is replaced by a call of a runtime service; the service inspects the current thread and performs on-stack replacement. Two approaches differ in only how to schedule threads requiring on-stack replacement, but share most of other code. The OSR runtime service can be broken down to several steps:

1. extract execution state from the top stack frame of a suspended thread,

2. generate new code for the suspended activation, and

3. transfer execution in the suspended thread to new compiled code.

Bytecode instructions for setting up new stack frames must be selected according to the types of values. Type information in OSR maps serves this purpose. One challenge of OSR transition is how to carry reference type values from old stack frames to new stack frames.

Values extracted from stack frames are in binary format. If a value is a reference, it is indeed the address of a live object. It is not feasible to use the row addresses during the OSR transition because the process may trigger GC and objects get moved. Jikes RVM provides primitives to disable and enable GC. If a system method needs to manipulate row addresses of some objects, it should disable GC first. After manipulation, it re-enables GC. We call such a region GC-critical. A GC-critical region has to be small, and the code in the region should not trigger GC. OSR transition code does not fit in one GC-critical region.

To solve this problem, we only create a GC-critical region for converting row addresses to normal Java references when extracting *JVM scope descriptors*. These references are saved in a temporary object array, which is subject to normal garbage collection. In the prologue of specialized method, loading a reference value to stack or local variable is done by bytecode instructions reading the element of the object array (indexes are known when storing references into the array).

Based on the implementation, Fink conducted some experiments of deferred compilation and long-running loop promotion. More experimental results can be found in the paper [FQ03]. On-stack replacement is considered as a powerful, yet expensive, invalidation mechanism. Therefore, an application has to use it wisely. Ideally, it should be used in the situation where invalidation is rare. We have not used on-stack replacement to perform direct inlining.

4.5 Related work and discussion

Guarded inlining uses runtime tests to ensure the correctness of optimized program. The optimized code preserves the semantics of original program regardless the input data and execution environments. However, code patching, preexistance based inlining, and on-stack replacement present a new scenario for optimizations, where optimized code is correct only with respect to the execution environment at optimization time. These techniques ensure that the system has ability to correct the invalidated code if future execution violates optimization assumptions. However, if the assumption is most likely to be true in the future, speculation enables more effective and aggressive optimizations. One common of three techniques is that the system knows when a speculation is violated and takes safety measurements before the execution goes wrong.

One application of on-stack replacement is to choose compilation and optimization units better than methods in a JIT environment. Whaley's partial method compilation [Wha01] essentially applies SELF-91's uncommon branch extension to "rare" blocks as determined by heuristics or profile data. His work assumes on-stack replacement is ready for use. Suganuma et. al [SYN03b] extends Whaley's approach further to use regions as compilation units. A region is a collection of code from several methods excluding rarely executed portions. At region exit point, it requires on-stack-replacement to transfer the execution to the original methods.

Bruening and Duesterwald [BD00] explored alternative compilation units for Java. Their results suggest that using hot loops and traces, in combination with methods, as compilation units can reduce compiled code size while preserving acceptable coverage of optimized code.

The techniques we discussed in this chapter ensure the correctness of speculative optimizations. The execution of optimized code preserves the semantics of original program at anytime. A more aggressive approach is to speculatively execute the program and correct the machine state if the speculation is wrong. RePLay [PL01] is a proposed hardware framework supporting speculative execution. We are not aware of software approaches that support speculative execution.

Chapter 5
Online Call Graph Construction

In this chapter, we study the call graph construction problem in Java virtual machines. The chapter is organized as follows. First we introduce the problem and motivation in Section 5.1. Section 5.2 reviews several classical static call graph construction algorithms for object-oriented programs. Then we describe three runtime type analyses for computing conservative call graphs in Section 5.3. A new call graph profiling mechanism is introduced in Section 5.4. We evaluated each algorithm by analyzing the cost of the analysis and comparing the quality of constructed call graphs on a set of standard Java benchmarks. Finally the related work and conclusion is discussed in Section 5.5.

5.1 Motivation

*Inter*procedural analyses (IPAs) derive more precise program information than *intra*procedural ones. Static IPAs provide a conservative approximation of runtime information to clients for optimizations. A foundation of IPA is the call graph of the analyzed program.

A call graph is a directed graph that represents call relations between methods (or functions in the C programming language). There exists a directed edge from a method A to a method B if A calls B. The precision of a call graph can be measured by two metrics: flow-sensitivity and context-sensitivity. A *flow-sensitive* call graph differentiates edges from different call sites in the same caller to a callee. For example, the method A may have two call sites s_1 and s_2 both calling the method B, a flow-insensitive call graph has one edge from A to B, but a flow-sensitive graph has two distinct edges from $A.s_1$ to B and $A.s_2$ to B. Flow-sensitive call graphs are useful for flow-sensitive analyses, and flow-insensitive call graphs are usually more compact.

The second measurement of call graphs is the context-sensitivity. Figure 5.1(a) is a simple example for showing call graphs with different context-sensitivity. Figure 5.1(b) shows the call graph without contexts (we call it 0-degree context-sensitive). This is not very useful because an analysis has to assume that a method can be called by all methods. A most commonly used call graph is the one with 1-degree context-sensitivity, as shown in Figure 5.1(c), where each method is represented as one node in the graph and edges are call relations between methods. Usually we call it *context-insensitive* comparing to ones with higher degrees shown in Figure 5.1(d). In Figure 5.1(d), a callee has one representative node for each calling context. In this study, we focus on constructing *context-insensitive* call graphs since it is most widely used by interprocedural analyses.

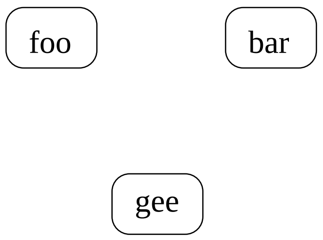
A main feature of object-oriented (OO) programming languages is the support of polymorphism. A polymorphic call is a call site which may invoke different method during program execution. Polymorphism has big engineering benefits for program design, code reuse and easy maintainance, and it is typically used for developing large frameworks, the Java utility library is such an example. In the Java programming language, polymorphism is implemented as virtual method calls, which target is looked up by the type of receiver object and callee signature at runtime. Although virtual calls increase the flexibility of program design, they incur large runtime overhead comparing to static calls whose targets can be statically binded at compile time. From compiler writers, virtual calls also pose difficulties on program analyses and optimizations.

Since a program analysis analyzes the code before its execution, it is not possible to know the exact runtime types of a receiver at a virtual call site. However, language constraints and program contexts can let the analysis compute a super set of the receiver's runtime types. The central problem of static call graph construction for

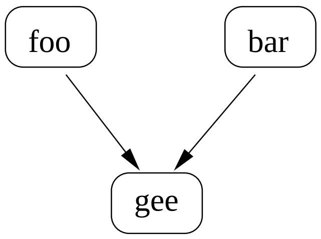
```
foo() {    bar() {
 gee();      gee();
}          }

gee() {...}
```

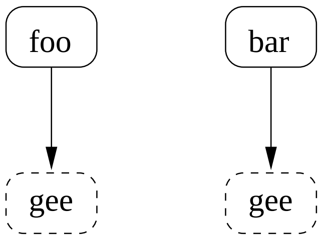
(a) a simple example



(b) call graph without context (0-degree)



(c) context-insensitive call graph (1-degree)



(d) context-sensitive call graph (2-degree)

Figure 5.1: Call graphs with different context-sensitivity

object-oriented programs is often converted to efficient computation of accurate type sets of receiver variables at virtual call sites.

In addition to virtual calls, call graph construction for Java is further complicated by the presence of dynamic class loading. Static call graph constructors assume that all code that may be executed at runtime is available for analyzing. In a Java virtual machine, however, classes are loaded dynamically to reduce resource usage. It is also beneficial to delay the resolution of symbolic references as late as possible until the referred entity is required. A call graph constructor or program analysis should avoid triggering the resolution of these unresolved references. Therefore, a dynamic call graph has to be incremental (dealing with dynamic class loading), efficient, and type safe.

In next several sections, we show a general approach for handling Java's dynamic features seamlessly in a JIT environment. Our approach also has little overhead on the execution of applications. First, we review classical static call graph construction algorithms for object-oriented programs. Then we show how to adapt and extend several static type analyses to runtime for computing dynamic call graphs. We introduce a new mechanism that uses a profiling code stub to capture invoked call edges. The call graph constructed by our profiling mechanism is dramatically smaller than those computed by type analyses. All algorithms are designed to be efficient and can be used in practice. A very desirable feature of our approach is that call graphs can be built incrementally while execution proceeds. An interprocedural analysis based on dynamic call graphs can support speculative optimizations.

5.2 Static call graph construction for OO programs

A static call graph is constructed by using a type analysis to compute runtime type sets of reference variables. The type set is then used, together with a callee signature, to resolve the set of targets of a virtual call. A static type analysis requires all classes are available and a complete class hierarchy can be constructed at analysis time.

5.2.1 Class hierarchy analysis and rapid type analysis

In a strongly typed language such as Java, each variable has a declaring type. At runtime, the variable can only point to objects of its declaring type and subtypes. Class hierarchy analysis (CHA) [DGC95] exploits this language constraint. CHA makes a conservative assumption that all subtypes of a receiver's declaring type are possible types at runtime (the most conservative assumption is that all types are possible, and this yields no useful information for program analyses).

Given a complete class hierarchy and a type T, we define a set, `hierarchy_types(T)`, be the type T and its subtypes in the hierarchy. Given a call site s whose receiver's declaring type is C, the class hierarchy analysis assumes *hierarchy_types(C)* is the runtime type set of the receiver (in other words, all types not in *hierarchy_types(C)* cannot be receiver types of s). For Java programs, we can limit *hierarchy_types* to only *normal* classes (vs. interface) since interfaces cannot be instantiated. *Abstract* classes are regarded as normal too. If C is an interface, *hierarchy_types(C)* includes normal classes that implement C directly or indirectly. The cost of static CHA is merely the cost of constructing class hierarchy, which can be done very easily and quickly by any modern compiler.

Rapid type analysis (RTA) [BS96] uses program contexts to refine type sets computed by CHA. Given a program P, only types with allocation sites in P's program text can be instantiated at runtime. Therefore, the type set of a receiver can be pruned by removing classes in *hierarchy_types(C)* that do not have allocation sites in the program P.

To gather types with allocation sites, RTA needs to parse the program text linearly. The time and space required by RTA is minimal. RTA is considered to be a fast and effective improvement of CHA.

5.2.2 Reachability-based interprocedural type analysis

Further constraints can be derived from program texts by analyzing and tracking data flows. Reachability-based algorithms [Ste96, And94, TP00, SHR+00] build flow graphs for the program at different granularity levels. Variables have representative nodes in the flow graph, and assignments are formulated as flow edges between nodes. Interprocedural data-flow is built on a call graph computed by CHA or RTA.

If an allocation site can reach a variable in the flow graph, the allocated type is considered to be in the runtime type set of the variable. A variable's declaring type can be used to filter out reachable types that are not in its subtypes, either during or after the propagation.

Algorithms often trade the time and space with precision. Context-insensitive algorithms can be modelled as unification-based [Ste96] or subset-based [And94] propagation as points-to analysis. The complexity varies from $O(N\alpha(N,N))$ for unification-based analysis to $O(N^3)$ for subset-based analysis. Context-sensitive algorithms [EGH94, WL95] might yield more precise results but are difficult to scale to large programs.

The results of a static type analysis can be used to prune the basic call graph or guide the inlining of virtual calls. In this thesis, we developed online versions of two reachability-based type analyses, XTA [TP00] and VTA [SHR+00]. Since these type analyses are performed in a JIT compiler, the results are only used for method inlining in our study.

5.3 Runtime call graph construction using type analyses

Now we discuss how to adapt static CHA and RTA to dynamic ones for incrementally constructing call graphs in a JIT compiler. We present a general approach for handling dynamic class loading and lazy resolution of symbolic references. We also describe a new dynamic type analysis, ITA, which uses unique runtime allocation information to further improve the results of CHA and RTA.

5.3.1 Dynamic CHA and RTA for call graph construction

Dynamic CHA. At runtime, a Java virtual machine maintains a hierarchy of loaded classes. As program's execution proceeds, new classes may be loaded and added to the hierarchy. Thus, the hierarchy tree grows dynamically. The compiler can compute a dynamic *hierarchy_types* set of C by walking through the snapshot of the hierarchy at an execution point. In contrast to the static *hierarchy_types(C)*, the dynamic one can expand as the class hierarchy grows.

In a Java virtual machine, a class has to be initialized before its first instance gets created (see Section 1.1 for more details on class loading process). Given a program P, we define the set of initialized classes as *initialized_types(P)*. Since class initialization is triggered by the virtual machine, the set *initialized_types(P)* is built and dynamically expanded by the runtime system. Given a variable o with a declaring type T, the dynamic CHA computes the type set of o by

$$type\_sets(o) = hierarchy\_types(T) \cap initialized\_types(P)$$

Dynamic RTA. In Java virtual machines, a method is parsed by interpreters or JIT compilers before its execution. Dynamic RTA can collect all allocation sites of parsed methods encountered so far.

Given a program P, we define the set of types used by allocation expressions in parsed methods of P as *rapid_types(P)*[1]. The type set of o is defined as the intersection of *hierarchy_types(T)* and *rapid_types(P)*:

$$type\_sets(o) = hierarchy\_types(T) \cap rapid\_types(P)$$

The *rapid_types(P)* contains only initialized classes (initialized implies resolved), therefore, it is strictly smaller than *initialized_types(P)*.

To unify implementations of dynamic CHA and RTA, we define a meta type set *eligible_types(P)* which is *initialized_types(P)* when using CHA or *rapid_types(P)* when using RTA. Because the bytecode does not carry the static declaring type of

[1]Classes that create objects via *newInstance* method are treated as members of *rapid_types(P)* immediately.

a variable from the source language, we use the declaring type of the callee method for computing call targets. Java's static type system implies this is a safe solution. When analyzing a call site, S, with a resolved method reference of $C.m$, The analysis computes possible targets using the following algorithm:

```
for each class sC in hierarchy_types(C)
  if sC is not in eligible_types(P)
    continue;
  while sC != null
    if sC declares m
      generate a call edge from S to sC.m
      break;
    else
      sC = sC's super class
```

One caveat of dynamic RTA is that when the JIT compiler compiles an allocation site with a type reference `mref`, the reference may not be resolved yet. An unresolved symbolic reference is only the name of a class. Now we discuss handling of dynamic class loading and unresolved type references.

Dealing with dynamic class loading. In a Java virtual machine, dynamic class loading can happen in two forms: by implicitly accessing a class or a class member, or by explicitly load a class using `Class.forName` method. Dynamic class loading causes the hierarchy tree to grow, and expands *hierarchy_types* sets. When using dynamic CHA, the *initialized_types(P)* set is dynamically expanded as well. As for dynamic RTA, new members are added into *rapid_types(P)* when new methods are parsed. All that means analyzed call sites may have new runtime types, and both analyses must fix them properly by considering the expansion of type sets.

Our solution to type set expansion is to maintain a map $resolved\_sites : m \rightarrow \{s, \ldots\}$, where m is a *resolved* method, and s is a parsed call site whose resolved callee signature is m. Let T be a new member of the meta type set *eligible_types(P)*, dynamic CHA and RTA fix parsed call sites using the following algorithm:

```
for each virtual method m declared in T
for each method m' overridden by m
  for each s of resolved_sites(m')
    generate a call edge from s to m
```

Handling unresolved references. At compile time, a call site may have an unresolved *method reference* which cannot be resolved without loading new classes. A call graph constructor should not resolve such references because class loading incurs runtime overhead. Instead, unresolved call sites are put in a separate database using the method reference as index: $unresolved\_sites : mref \rightarrow \{s, \ldots\}$. The database monitors method reference resolution events. When a method reference *mref* is resolved to a method *m*, a new entry $m \rightarrow unresolved\_sites(mref)$ is added to the $resolved\_sites$ database discussed in previous paragraph.

Unresolved *type references* must be dealt with correctly for dynamic RTA. The analysis may encounter allocation sites whose allocation types are unresolved. An unresolved type reference is only a class name. To ensure the correctness, classes resolved from these references should be added into the *rapid_types(P)* as soon as they are initialized. Our solution is to track the set of unresolved types from parsed allocation sites, and monitors type reference resolution events. A newly resolved type from the set is marked and added into *rapid_types(P)* after its initialization.

Dynamic CHA and RTA do not need runtime checks in application code. All events triggering the analyses happen at class loading, reference resolution and method compilation time. The only cost is to maintain dependency databases that incur memory overhead. In our implementation, we use special integer sets and hash tables for saving memory.

5.3.2 Instantiation-based type analysis

Dynamic RTA improves the results of CHA by reducing the size of the meta set *eligible_types(P)* of a program *P*. Java explicitly requires garbage collection as its dynamic memory management. Objects are allocated through a memory management interface. This allows the virtual machine know the exact set of classes that have instances at runtime. Only classes with instances can be the runtime types of a receiver. This leads to a new improvement of the class hierarchy analysis, `instantiation-based type analysis` (ITA).

During the execution of a program P, we define a set *instantiated_types(P)* as the set of types that have had instances. It is easy to see that *instantiated_types(P)* is a subset of *rapid_types(P)*[2]. Given a variable o with a declaring class C, the type set computed by ITA is the intersection of *hierarchy_types(C)* and *instantiated_types(P)*:

$$\mathit{type\_set}(o) = \mathit{hierarchy\_types}(C) \cap \mathit{instantiated\_types}(P)$$

In Jikes RVM, objects are created in heaps via allocators. An allocator takes a type and returns an object. Building *instantiated_types(P)* is straightforward. The allocation sequence checks if the type passed in was in the set. If not, it adds the type as the new member to the set. When building the bootimage, the bootimage writer builds the type set by scanning objects in the bootimage. The meta type set, *eligible_types(P)*, is *instantiated_types(P)* for ITA, and algorithms for dealing with unresolved method references and type set expansion are shared among CHA, RTA and ITA.

When compilers compile a *new* bytecode instruction, if the type reference is resolved and is in the *instantiated_types(P)* set, then the instruction is compiled to a runtime service without checks. This can happen in two situations: another allocation site of the same type was executed before compiling the current site; the current allocation site is being recompiled in an adaptive system.

Table 5.1 shows the number of scalar allocations that require a check at runtime in our benchmarks. The second column shows the dynamic counts of total allocations during benchmark runs. The third column lists the number of allocations requiring checks, followed by the percentages of total allocation counts. Except *_201_compress* and *_222_mpegaudio*, which are not allocation intensive, other benchmarks only need checks for a small portion of allocations. Overall, adding a check in the allocation sequence does not cause measurable effects.

[2]Classes that create objects via *newInstance* method are treated as members of *rapid_types(P)* and *instantiated_types(P)*.

benchmark	total	checks	
_201_compress	455,488	450,585	(99%)
_202_jess	57,410,199	3,760,777	(7%)
_209_db	32,247,049	1,036,771	(3%)
_213_javac	54,486,512	20,475,862	(38%)
_222_mpegaudio	1,817,243	1,789,324	(98%)
_227_mtrt	61,766,963	4,643,998	(8%)
_228_jack	51,622,992	10,318,194	(20%)
SpecJBB2000	347,774,820	119,048,760	(34%)
CFS	8,085,943	2,726,610	(34%)

Table 5.1: The number of scalar allocations with checks

5.3.3 Characteristics of dynamic CHA, RTA, and ITA

Dynamic ITA and RTA improves CHA by reducing the size of the meta set *eligible_types(P)*. Table 5.2 shows the size of the meta type set at the end of benchmark runs. Columns 2 to 4 show the total numbers of classes in *eligible_types(P)* used by CHA, RTA and ITA. The rest of columns break the number further down to two categories: classes belonging to Jikes RVM and classes in applications and libraries[3]. The numbers of classes from Jikes RVM are very close for different benchmarks, the numbers of application and library classes vary a lot. The size difference between *initialized_types(P)* and *rapid_types(P)* is pretty small in all categories. It indicates that the dynamic RTA is less effective on improving CHA. However, among initialized classes, only less than half of them have created instances at runtime. This can be explained that many kinds of bytecode instructions can trigger class loading and initialization. For example, static field accesses and static method invocations can cause class loading and initialization, and if a class is initialized, its superclasses must be initialized too. The much smaller *instantiated_types(P)* indicates ITA could improve the results of CHA more than dynamic RTA.

[3]We use the package name to distinguish the source of classes, some library classes are used by both the RVM and the applications.

	all			RVM			app&lib		
benchmark (P)	CHA	RTA	ITA	CHA	RTA	ITA	CHA	RTA	ITA
_201_compress	1160	1042	541	854	767	409	306	275	132
_202_jess	1290	1169	672	854	767	414	436	402	258
_205_raytrace	1173	1053	557	854	767	413	319	286	144
_209_db	1154	1035	542	854	767	416	300	268	127
_213_javac	1286	1163	658	854	767	419	443	396	239
_222_mpegaudio	1190	1063	567	854	767	414	336	296	153
_227_mtrt	1173	1053	557	854	767	413	319	286	144
_228_jack	1194	1073	575	854	767	413	340	306	162
SpecJBB2000	1264	1130	643	854	767	415	410	363	228
CFS	1185	1060	571	854	767	415	331	293	156

Table 5.2: Statistics of *initialized_types(P)*, *rapid_types(P)* and *instantiated_types(P)*

5.3.4 Evaluation

In this section we evaluate dynamic CHA, RTA and ITA for constructing call graphs. Jikes RVM is written in Java and it has many more classes than any benchmark we have. If an analysis includes RVM classes, the results would be overwhelmed by the information from these classes. The production configuration of Jikes RVM is a *FastAdaptive* setup (it can choose different garbage collectors). *FastAdaptive* configuration pre-compiles all RVM classes (the runtime system, compilers, garbage collectors, and some core Java libraries) into the bootimage (native code and resources). Although this configuration takes longer time to build the bootimage and may have larger memory footprint, the execution is much faster than a thin configuration that has to compile the optimizing compiler itself at runtime. In our evaluation of call graph construction, we use the *FastAdaptive* configuration.

To construct call graphs for bootimage classes, dynamic CHA, RTA and ITA require the bootimage compiler to collect call sites of methods compiled into the bootimage. All call sites are treated as unresolved and registered in *unresolved_sites(P)*.

	ALL		non-virtual	virtual		interface	
CHA	146,740		35,292	51,706		59,742	
RTA	145,427	99.1%	35,292	50,660	98.0%	59,475	99.6%
ITA	69,893	47.6%	35,292	25,898	50.0%	8,703	14.6%

Table 5.3: Call graph size of bootimage classes only

Besides the registered call sites, dynamic CHA does not require special care because the class hierarchy is built in the bootimage. Dynamic RTA requires the compiler to mark types of allocation sites as in *rapid_types(P)*. When copying objects from host Java virtual machine to the RVM bootimage, ITA puts object types into *instantiated_types(P)*. All classes in the bootimage are initialized.

When executing an application, Jikes RVM first loads the bootimage into the heap and initializing other necessary resources. Before starting the main application thread, the RVM calls the type analysis which goes through registered call sites from the bootimage and builds call edges for resolved ones just as they are newly compiled.

Table 5.3 shows the number of call edges built by dynamic CHA, RTA, and ITA, for bootimage classes only. All call graphs are *flow-sensitive* unless otherwise stated. The second column shows the total number of call edges in the graph, and columns 3 to 5 breaks them further down to three categories according to the type of each call sites. *Non-virtual* calls includes call edges from *invokestatic* and *invokespecial*. Since these *non-virtual* calls are treated in the same way by analyses, the number of call edges are the same as well. We are interested in the last two columns: call edges from *invokevirtual* (column 4) and *invokeinterface* (column 5). Comparing to the call graph built by CHA, dynamic RTA has nearly no improvement of call graphs constructed by CHA. Dynamic ITA removes about 58% edges from *invokevirtual* and 86% from *invokeinterface*. However, this improvement is not meaningful since the bootimage is statically compiled. It only serves the purpose to be compared with Table 5.4.

Table 5.4 compares the call graph sizes in the same way as Table 5.3 except that the call graph contains application and library classes of *_213_java* benchmark. When

	ALL		non-virtual	virtual		interface	
CHA	167,519		37,297	65,779		64,443	
RTA	166,053	99.1%	37,297	64,580	98.1%	64,176	99.6%
ITA	131,206	78.3%	37,297	57,440	87.3%	36,469	56.6%

Table 5.4: (_213_javac) call graph size when including bootimage classes

application and library classes participate the analysis, dynamic ITA removes 13% of call edges from *invokevirtual* and 44% from *invokeinterface*. Recall that dynamic ITA only has negligible runtime overhead, it is a good replacement of CHA for building a call graph including the bootimage.

Now we look at call graph sizes for application and library classes only. Table 5.5 compares edges numbers by dynamic CHA, RTA, and ITA of all benchmarks. Columns have the same meaning as Table 5.3 and 5.4. *Invokevirtual* bytecode contributes more edges than other kinds of calls. Using CHA constructed call graph as the base, dynamic RTA removes only 2 to 3% edges from *invokevirtual*, and ITA reduces the number of edges up to 10%. Dynamic ITA is less effective on application classes than on bootimage classes.

From Table 5.3 and 5.5, we can see the dynamic CHA leaves a small room for improvement by other type analyses. This is a very different characteristic of static CHA and dynamic CHA.

It is worth to point out that the number of call edges of *_213_javac* in Table 5.4 is much larger than the sum of edge numbers from Table 5.3 and 5.5. The bootimage contains RVM classes that have been initialized. Methods other than static class initializers ($< clinit >$) are compiled but not executed. Many RVM classes only have instances when running applications. Thus, the *initialized_types(RVM)* from the bootimage does not contain many RVM classes which are used by compilers and garbage collectors at runtime.

Because RVM classes and application classes share the same class hierarchy and the same meta type set, mixing them together generates false call edges from RVM call sites to application methods and from application call sites to RVM methods too.

benchmark		ALL	non-virtual	virtual		interface
_201_compress	CHA	1862	1057	774		31
	RTA	1843	1057	755	97.5%	31
	ITA	1802	2057	730	94.3%	15
_202_jess	CHA	4325	1762	2070		493
	RTA	4286	1762	2033	98.2%	491
	ITA	4171	1762	1955	94.4%	454
_205_raytrace	CHA	3023	1241	1751		31
	RTA	3003	1241	1731	98.9%	31
	ITA	2951	1241	1695	96.8%	15
_209_db	CHA	2177	1107	941		129
	RTA	2157	1107	921	97.9%	129
	ITA	2059	1107	895	95.1%	57
_213_javac	CHA	14936	2038	12390		508
	RTA	14783	2038	12227	98.7%	508
	ITA	13796	2038	11502	92.8%	256
_222_mpegaudio	CHA	2520	1236	1239		45
	RTA	2470	1236	1189	96.0%	45
	ITA	2412	1236	1147	92.6%	29
_227_mtrt	CHA	3024	1242	1751		31
	RTA	3004	1242	1731	98.9%	31
	ITA	2952	1242	1695	96.8%	15
_228_jack	CHA	4658	1940	2278		440
	RTA	4618	1940	2238	98.2%	440
	ITA	4352	1940	2204	96.8%	208
SpecJBB2000	CHA	7186	2002	4775		389
	RTA	7128	2002	4717	98.8%	389
	ITA	6923	2002	4623	96.8%	278
CFS	CHA	3385	689	2545		151
	RTA	3331	689	2491	97.9%	151
	ITA	3074	689	2285	90.0%	100

Table 5.5: Call graph size comparison of dynamic CHA, RTA, and ITA (application and libraries)

In order to get effective optimizations on application code in Jikes RVM, we need a sound approach to distinguish RVM classes from application classes instead of just using package names. Unfortunately, this is currently an open question in research community using Jikes RVM.

5.4 Accurate call graph construction using profiling code stubs

Dynamic type analysis, such as CHA, RTA, and ITA, can be used to build conservative call graphs at runtime. However, it is desirable to have a more precise call graph for most interprocedural analyses. Also we would like to find the answer of the question: how precise are these call graphs built by type analyses?

Instead of using type analysis to compute a conservative call graph, we propose a dynamic approach for profiling and constructing context-insensitive call graphs at runtime. The mechanism initializes virtual method tables using a profiling code stub. When a method call happens, the code stub is executed. The code stub generates a call edge event, then triggers method compilation if the method is not compiled yet, and finally patches the virtual method table using the method code address.

The proposed mechanism has some very desirable features as a runtime call graph constructor:

Accuracy. Only executed call edges are in the call graph.

Efficiency. It captures the first invocation event of each call edge, and only the first execution has some profiling overhead. The repeated calls only need to execute at most one more instruction. We also show several optimizations to reduce the overhead further.

Just-in-time: A call edge is captured before the control flow is transferred to the callee. Clients, such as call graph builders, can register callback routines called by the profiling code stub when new call edges are discovered. Callbacks can perform necessary actions before the callee is invoked. This enables speculative optimizations with invalidation.

Flexibility: The profiling mechanism can be used in conjunction with type analyses. For different kinds of call sites, we can choose profiling or type analysis for trading off efficiency and accuracy.

The remainder of this section is structured as follows. First, in Section 5.4.1 we give the necessary background, describing the implementation of virtual method tables in Jikes RVM. In Section 5.4.2 we describe the basic mechanism we propose for building call graphs at runtime, and in Section 5.4.3 we show how this basic mechanism can be optimized to reduce overheads. Finally, in Section 5.4.4, we measure the cost of call graph profiling and compare profiled call graphs to ones constructed by using dynamic CHA.

5.4.1 Background: virtual method table

Before jumping into details of the profiling mechanism, it is necessary to understand how virtual method calls are implemented in Jikes RVM. We first revisit the virtual method dispatch table in Jikes RVM [AAB+00], which is a standard implementation in modern Java virtual machines. Figure 5.2 depicts the object layout in Jikes RVM. Each object has a pointer, in its header, to the Type Information Block (TIB) of its type (class). A TIB is an array of objects that encodes the type information of a class. At a fixed offset from the TIB header is the Virtual Method Table (VMT) which is embedded in the TIB array. A resolved method has an entry in the VMT of its declaring class, and the entry offset to the TIB header is a constant, say `method_offset`, assigned during class resolution. A VMT entry records the instruction address of the method that owns it. Figure 5.3 shows that, if a class, say `A`, inherits a method from its superclass, `java.lang.Object`, the entry at the method offset in the subclass' TIB has the inherited method's instruction address. If a method in the subclass, say `D`, overrides a method from the superclass, the two methods still have the same offset, but the entries in two TIBs point to different method instructions.

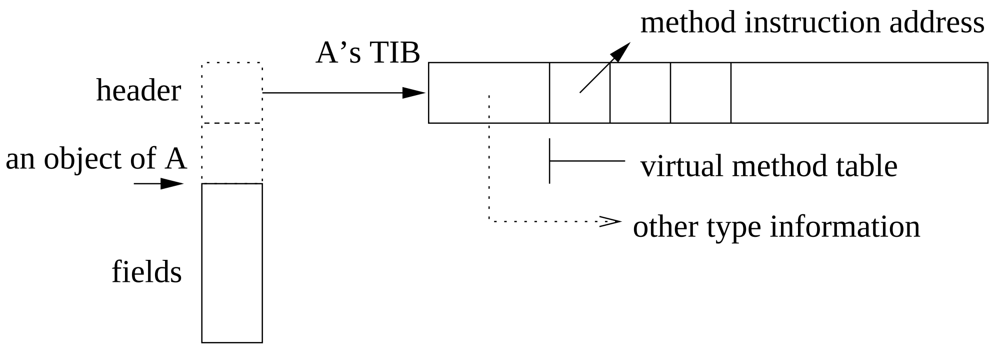


Figure 5.2: TIB in Jikes RVM

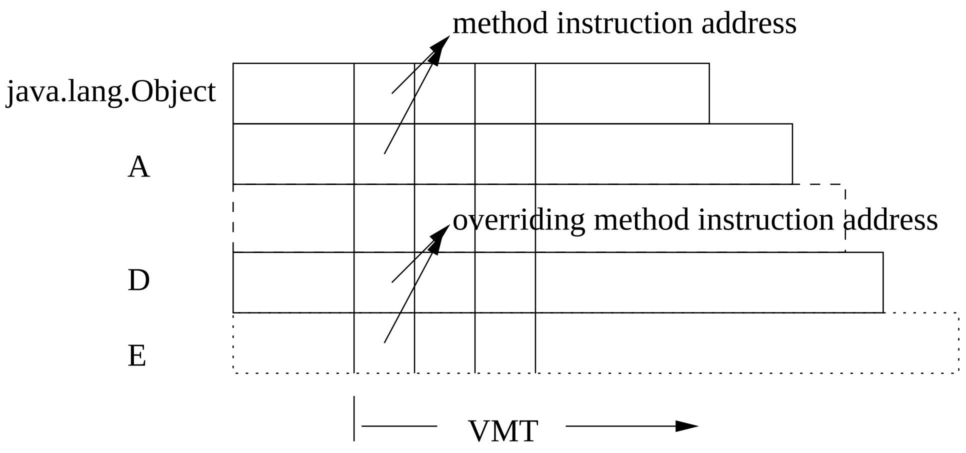

method instruction address
java.lang.Object
A
overriding method instruction address
D
E
VMT

Figure 5.3: VMT in Jikes RVM

Given an object pointer at runtime, an *invokevirtual* bytecode is implemented by three basic operations:

```
TIB = * (ptr + TIB_OFFSET);
INSTR = TIB[method_offset];
JMP INSTR
```

The first instruction obtains the TIB address from the object header. The address of the real target is loaded at the `method_offset` offset in the TIB. Finally the execution is transferred to the target address.

Lazy method compilation works by first initializing TIB entries with the address of a lazy compilation code stub. When a method is invoked for the first time, the code stub gets executed. The code stub triggers the compilation of the target method and patches the address of the compiled method into the TIB entry (where the code stub resided before).

5.4.2 Call graph construction by profiling

Lazy method compilation code stub captures the first invocation of a method without distinguishing callers. It can be viewed as constructing a call graph with 0-degree context-sensitivity (e.g., Figure 5.1(b)) where a method can be called by all compiled methods. The space overhead of lazy method compilation is that each method requires a TIB entry. If there are n methods, it requires n TIB slots.

Given n methods, there are n^2 possible call edges in an 1-degree context-sensitive call graph (e.g., Figure 5.1(c)). In order to capture call edges, we extended the TIB structure to store information per caller. Figure 5.4 shows the extended TIB structure. The TIB entry of a method is replaced by an array of instruction addresses. We call the array a Caller-Target Block (CTB). The indexes of CTB slots (*caller_index*) are dynamically assigned to callers[4] of the method by the JIT compilers. The up-bound of memory overhead for CTBs could be n^2. But dynamic assignment of CTB indexes reduces the memory requirement to a very low level. Note that now an *invokevirtual* bytecode takes one extra load to get the target address.

[4]or call sites if the call graph is flow-sensitive.

```
TIB = * (ptr + TIB_OFFSET);
CTB = TIB[method_offset];      /* load method's CTB array from TIB */
INSTR = CTB[caller_index];     /* load method's code address */
JMP INSTR
```

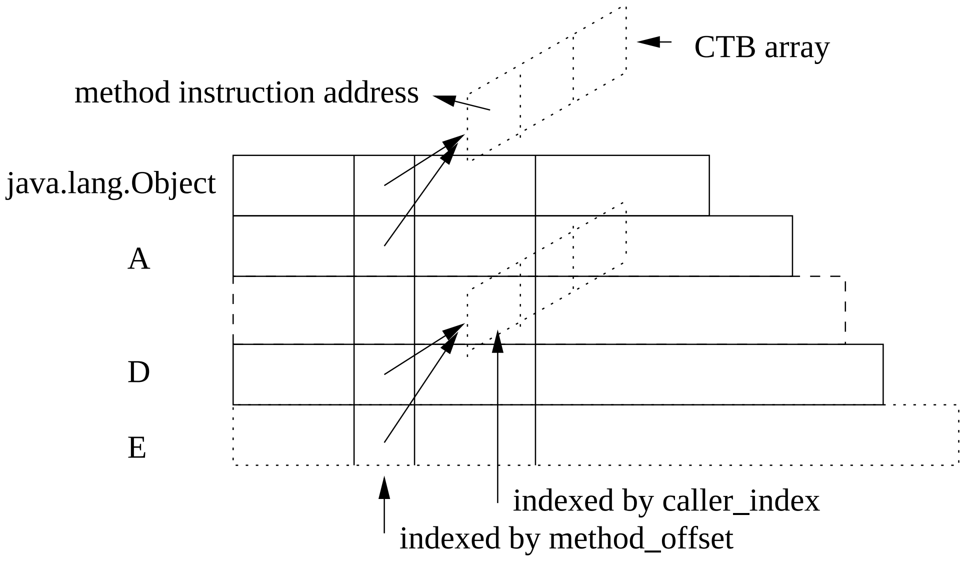


Figure 5.4: Extended VMT for profiling call graph

The lazy method compilation code stub is extended to a profiling code stub which, in addition to triggering the lazy compilation of the callee, also generates a new call edge event from the caller to the callee. Initially all of the CTB entries have the address of the profiling code stub. When the code stub at a CTB entry gets executed, it notifies clients monitoring new call edge events, and compiles the callee method if necessary. Finally the code stub patches the callee's instruction address into the CTB entry. Clearly the profiling code stub at each entry of the CTB array will execute at most once, and the rest of the invocations from the same caller will execute the callee's machine instruction directly.

There remain four problems to address. First, one needs a convenient way of indexing into the CTBs which works even in the presence of unresolved method references and virtual calls. Second, the implementation of interface calls should be aware of the CTB array. Third, non-virtual calls (static methods and object initializers) can be handled specially. Fourth, we must handle the case where an optimizing compiler inlines one method into another. Our solution to these four problems is given below.

Allocating slots in the CTB

To index callers of a callee, our modified JIT compiler maintains a map from a callee method signature to an array of callers:

$$callercounter : callee \rightarrow callers[]$$

When the compiler compiles a virtual call to a *callee* in the *caller*, it checks whether $callercounter(callee)$ contains the *caller*. If *caller* is not in the map, it is put in *callee*'s caller array. The index of *caller* in the array is returned as the CTB index of the call site.

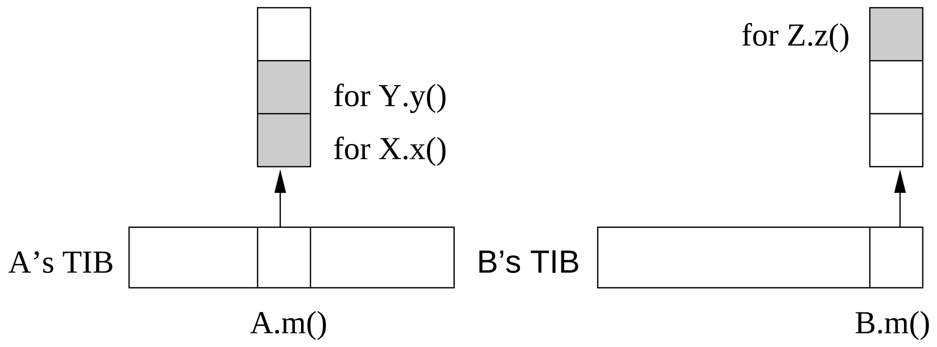


Figure 5.5: Example of allocating CTB indexes

In Java bytecode, an *invokevirtual* instruction contains only a symbolic reference to the name and descriptor of the method as well as a symbolic reference to the class where the method can be found. Resolving the method reference to a callee method signature requires the class to be loaded first. To deal with unresolved method references and virtual calls, our approach uses the callee's method name and descriptor as the index in the map instead of the resolved method:

$$callercounter : (name, desc) \rightarrow callers[]$$

For example, both methods `X.x()` and `Y.y()` have virtual calls of a symbolic reference `A.m()`, and another method `Z.z()` has a virtual call of `B.m()`. Because the references `A.m()` and `B.m()` may resolve to the same method at runtime, we take a conservative assumption that all three methods are possible callers of any method with the signature: $(m, ()) \rightarrow [X.x(), Y.y(), Z.z()]$[5], and allocates slots in the TIB for all of them. At runtime, only two CTB entries of `A.m()` may be filled, and only one entry of `B.m()` may get filled. Figure 5.5 shows what the CTBs look like for method `A.m()` and `B.m()`. With this solution no accuracy is lost, but some space may be wasted due to unfilled CTB entries. Although some space is sacrificed, our approach simplifies the task of handling symbolic references and virtual calls. In real applications we observed that only a few common method signatures, such as `equals(java.lang.Object)`, and `hashCode()`, have large caller sets where space is unused.

Approximating interface calls

Interface calls are considered to be more expensive than virtual calls in Java programs because a normal class can only have a single direct super class, but could implement multiple interfaces. Jikes RVM has an efficient implementation of interface calls using a interface method table with conflict resolution stubs [ACF+01].

We tried two approaches to handle interface calls in the presence of CTB arrays. Our first approach profiles interface calls by allocating a *caller_index* for a call site in the JIT compiler and generating an instruction before the call to save the index value in a known memory location. After a conflict resolution stub has found its target method, it loads the index value from the known memory location. The CTB array of the target method is loaded from the TIB array of receiver object's declaring class. The target address is read out from the CTB at the index, and finally the

[5]A full method descriptor should include the name of the method, parameter types, and the return type. In this example, we use the name and parameter types only for simplicity.

benchmark	ITA (s)	PROF (s)	
_201_compress	6.363	6.273	-1.1%
_202_jess	4.277	4.420	3.3%
_205_raytrace	2.650	2.745	3.6%
_209_db	12.635	12.722	0.6%
_213_javac	8.037	8.220	2.3%
_222_mpegaudio	5.422	5.629	3.8%
_227_mtrt	2.827	2.945	4.2%
_228_jack	4.831	4.930	2.0%

Table 5.6: Overhead of profiling interface calls (best of 10 runs)

resolution stub jumps to the target address. This approach uses two more instructions to store and load the *caller_index* than *invokevirtual* calls[6]. After introducing one of our optimizations in Section 5.4.3, inlining CTB elements into TIBs, the conflict resolution stub requires more instructions to check the range of the index value to determine if the indexed CTB element is inlined in the TIB or not. As shown in Table 5.6, the overhead of profiling interface call (with inlined CTB size of 4) ranges from -1.1% to 4.2% for SpecJVM98 benchmarks. Data were collected on a 1.5M Pentium M laptop with 512M memory, and benchmarks were run 10 times using `SpecApplication` driver with input size 100. We report the best run.

Our second approach was to simply use dynamic type analysis to compute call edges for *invokeinterface* call sites at compile time, without introducing profiling instructions.

Table 5.7 shows the number of call edges from *invokeinterface* calls using ITA type analysis and profiling. Although profiling (3rd column) reduces a large number of call edges, the absolute number of call edges from *invokeinterface* is only a small portion of total call edges. We chose to use the second approach for the remaining experiments in this thesis.

[6]Certainly, if there is a spare register for use, we can save the index in the register and read it out in the resolution stub, but registers are scarce resources in common architectures.

benchmark	ITA	PROF
_201_compress	15	7
_202_jess	454	144
_205_raytrace	15	7
_209_db	57	19
_213_javac	256	59
_222_mpegaudio	29	21
_227_mtrt	15	7
_228_jack	208	92
SpecJBB2000	278	37
CFS	100	18

Table 5.7: Call edges from *invokeinterface* by ITA and profiling

Handling static methods and object initializers

Because there are many object initializers that share a common name <`init`> and descriptor, their CTB arrays may grow too large if we allocate CTB slots using the name and descriptor as index. Since calls of object initializers and static methods are non-virtual, the allocation of CTB slots for each method is independent of other methods even with the same name and descriptor. For example, static methods `A.m()` and `B.m()` both can use the same CTB index for different callers. Therefore, there is no superfluous space in CTB arrays of object initializers and static methods. The only problem is to handle unresolved method references correctly. For these unresolved static or object initializer method references, a dependency on the reference from the caller is registered in a database. When the method reference gets resolved, the dependency is converted to a call edge conservatively. Table 5.8 shows the numbers of call edges constructed by ITA and profiling mechanism from static methods and object initializers, on our set of benchmarks. Using the *_213_javac* benchmark as example, ITA adds 87 more edges than profiling, but it is only about 1.5% more of total edges.

benchmark	static		init	
	ITA	PROF	ITA	PROF
_201_compress	179	157	201	147
_202_jess	368	332	632	560
_205_raytrace	215	192	320	272
_209_db	189	164	224	170
_213_javac	389	356	908	855
_222_mpegaudio	194	173	325	269
_227_mtrt	216	194	320	272
_228_jack	302	277	542	479
SpecJBB2000	788	726	1001	807
CFS	197	133	403	301

Table 5.8: Call edges for non-virtual calls by ITA and profiling

Dealing with Inlining

Optimizing compilers perform aggressive inlining on a few hot methods. We capture these events as follows. When a callee is inlined into a caller by an optimizing JIT compiler, the call edge from the caller to callee is added to the call graph unconditionally. This is a conservative solution without runtime overhead. Since an inlined call site is likely executed before its caller becomes hot, the number of added superfluous edges is modest. Table 5.9 validates our assumption. Column 2 shows the numbers of call edges when method inlining is disabled, and column 3 lists the edge numbers when inlining is enabled. The edge number increment ranges from 1.6 to 6.9% for most of our benchmarks except *_213_javac* and *CFS*. The last column shows the number of call edges created due to inlining events.

5.4.3 Optimizations

Our runtime call graph construction mechanism may incur two kinds of overhead in Jikes RVM. First, adding one instruction per call can potentially consume many CPU

benchmark	full	inlining		inlined	
_201_compress	1423	1446	1.6%	334	23%
_202_jess	3208	3430	6.9%	502	15%
_205_raytrace	2534	2585	2.0%	443	17%
_209_db	1588	1660	4.5%	363	22%
_213_javac	6012	6915	15.0%	1280	19%
_222_mpegaudio	1894	1940	2.4%	352	18%
_227_mtrt	2536	2587	2.0%	443	17%
_228_jack	3403	3524	3.6%	407	12%
SpecJBB2000	5214	5476	5.0%	528	10%
CFS	1611	1776	10.2%	467	26%

Table 5.9: Call edges due to inlined methods

cycles because Jikes RVM itself is compiled by the same compiler used for compiling applications, and it also inserts many system calls into applications for runtime checks, locks, object allocations, etc. Second, a CTB array is a normal Java array with a three-word header; thus CTB arrays can increase memory usage and create extra work for garbage collectors.

#callers	Java Libraries		_213_javac app	
0	2384	78.60%	325	27.71%
1	95	81.61%	167	41.94%
2-3	119	85.38%	120	52.17%
4-7	221	89.24%	185	67.95%
8-	339		376	
TOTAL	3159		1173	

Table 5.10: Distribution of CTB sizes (*_213_javac*)

Table 5.10 shows the distribution of the CTB sizes for *_213_javac* benchmark from SpecJVM98 suite [speb] profiled in a *FastAdaptive* bootimage. The bootimage contains mostly RVM classes and a few Java utility classes. We only profiled methods from Java libraries and the benchmark. A small number of methods from bootimage classes may have CTB arrays allocated at runtime because there is no clear cut mechanism for distinguishing between Jikes RVM code and application code. The first column shows the range of the number of callers. The second and third columns list the distributions of methods belonging to Java libraries and application code. To demonstrate that most methods have few callers, we calculated the cumulative percentages of methods that have no caller, ≤ 1, ≤ 3 and ≤ 7 callers in the first to fourth rows. We found that 89% of methods from (loaded classes in) Java libraries and 68% of methods from application code have no more than 7 callers. In these cases, it is not wise to create short CTB arrays because each array header takes 3 words. The last data row labelled "`TOTAL`" gives the total number of methods of all classes and the number of methods in each of two sub-categories.

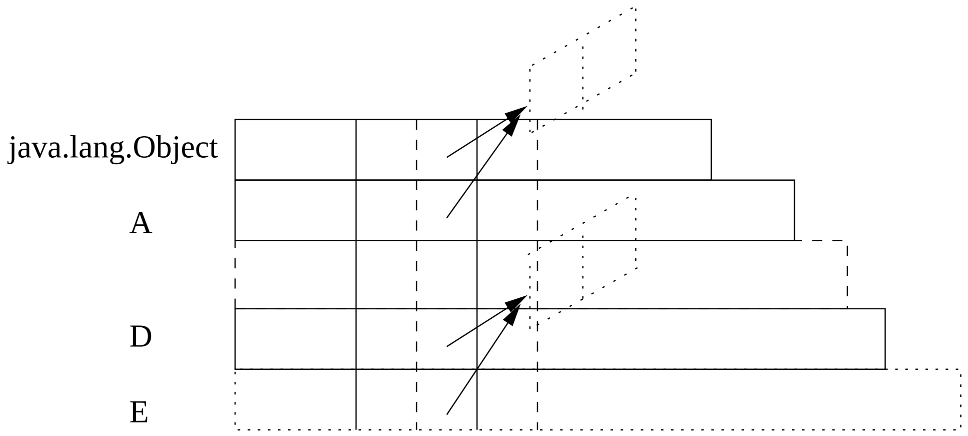


Figure 5.6: Inlining 1 element of CTB

To avoid the overhead of array headers for CTBs, and to eliminate the extra instruction to load the CTB array from a TIB in the code for *invokevirtual* instructions, a local optimization is to inline the first few elements of the CTB into the TIB. Since caller indexes are assigned at compile time, a compiler knows which part of the CTB will be accessed in the generated code. To accommodate the inlined part of the CTB, a class' TIB entry is expanded to allow a method to have several entries.

Figure 5.6 shows the layout of TIBs with one inlined CTB element. When generating instructions for a virtual call, the value of the caller's CTB index, `caller_index`, is examined: if the index falls into the inlined part of the CTB, then invocation is done by three instructions:

```
TIB = * (ptr + TIB_OFFSET);
INSTR = TIB[method_offset + caller_index];
JMP INSTR
```

Whenever a CTB index is greater than or equal to the inlined CTB size, *INLINED_CTB_SIZE*, then four instructions must be used for the call:

```
TIB = * (ptr + TIB_OFFSET);
CTB = TIB[method_offset + CTB_ARRAY_OFFSET];
INSTR = CTB[caller_index - INLINED_CTB_SIZE];
JMP INSTR
```

Note that in addition to saving the extra instruction for inlined CTB entries, the space overhead of the CTB header is eliminated in the common cases where all CTB entries are inlined.

Another source of optimization is to avoid the overhead of handling system code, such as runtime checks and locks, inserted by compilers, because this code is frequently called and ignoring them does not affect the semantics of applications. To achieve this, the first CTB entry is reserved for the purpose of system inserted calls. Instead of being initialized with the address of a call graph profiling stub, the first entry has the address of a lazy method compilation code stub or method instructions. When the compiler generates code for a system call, it always assigns the *zero* `caller_index` to the caller. To avoid the extra load instruction, the first entry of a CTB array is always inlined into the TIB.

5.4.4 Evaluation

We have implemented our proposed call graph construction mechanism in Jikes RVM [jikb] v2.3.0. Our benchmark set was introduced in Section 2.2. We use a variation of the *FastAdaptiveCopyMS* bootimage for evaluating our mechanism. In our experiment, classes whose names start with `com.ibm.JikesRVM` are not presented in the dynamic call graphs because (1) the number of RVM classes is much larger than the number of classes of applications and libraries, and (2) the classes in the boot image were statically compiled and optimized, type analysis such as ITA can be used to compute the call graph. Static IPAs such as extant analysis [Vug00] may be applied on the bootimage classes. We report the experimental results for application classes and Java library classes.

In our initial experiments we found that the default adaptive configuration gave significantly different behaviour when we introduced dynamic call graph construction because the compilation rates and speedup rates of compilers were affected by our call graph profiling mechanism. It was possible to retrain the adaptive system to work well with our call graph construction enabled, but it was difficult to distinguish performance differences due to changes in the adaptive behaviour from differences due to overhead from our call graph constructor. In order to provide comparable runs in our experiments, we used a counter-based recompilation strategy and disabled background recompilation. We also disabled adaptive inlining. This configuration is more deterministic between runs as compared to the default adaptive configuration. This behavior is confirmed by our observation that, between different runs, the number of methods compiled by each compiler is very stable. The experiment was conducted on a PC with a 1.5G Hz Pentium 4 CPU and 500M memory. The heap size of RVM was set to 400M. Note that Jikes RVM and applications share the same heap space at runtime.

The first column of Table 5.11 gives four configurations of different inlined CTB sizes and the default *FastAdaptiveCopyMS* configuration without the dynamic call graph builder. The bootimage size was increased about 10%, as shown in column 2, when including all compiled code for call graph construction. Inlining CTB elements increases the size of TIBs. However, changes are relatively small (the difference between inlined CTB sizes 1 and 2 is about 150 kilobytes), as shown in the second column.

The third column shows the memory overhead, in bytes, of allocated CTB arrays for methods of classes in Java libraries and benchmarks when running the *_213_javac* benchmark with an input size 100. The time for creating, expanding and updating CTB array is negligible.

Inlined CTB sizes	bootimage size (bytes)		CTB space (bytes)
default	24,382,332	N/A	N/A
1	26,809,420	9.95%	833,952
2	26,959,148	10.57%	814,104
4	27,218,672	11.63%	786,000
8	27,730,004	13.73%	746,944

Table 5.11: Bootimage sizes and allocated CTB sizes of *_213_javac*

A Jikes RVM-specific problem is that the RVM system and applications share the same heap space. Expanding TIBs and creating CTBs consumes heap space, leaving less space for the applications, and also adding more work for the garbage collectors. We examine the impact of CTB arrays on the GC. Since CTB arrays are likely to live for a long time, garbage collection can be directly affected. Using the *_213_javac* benchmark as example with the same experimental setting mentioned before, GC time was profiled and plotted in Figure 5.7 for the default system and configurations with different inlined CTB sizes. The x-axis is the garbage collection number during the benchmark run, and the y-axis is the time spent on each collection. We found that, with these CTB arrays, the GC is slightly slower than the default system, but not significantly. When inlining more CTB elements, the GC time is slightly increased. This might be because the increased size of TIBs exceeds the savings on CTB array headers when the inlining size gets larger. We expect a VM with a specific system heap would solve this problem.

The problem mentioned above also poses a challenge for measuring the overhead of call graph profiling. Furthermore, the call graph profiler and data structures are written in Java, which implies execution overhead and memory consumption, affecting benchmark execution times. To only measure just the overhead of executing profiling

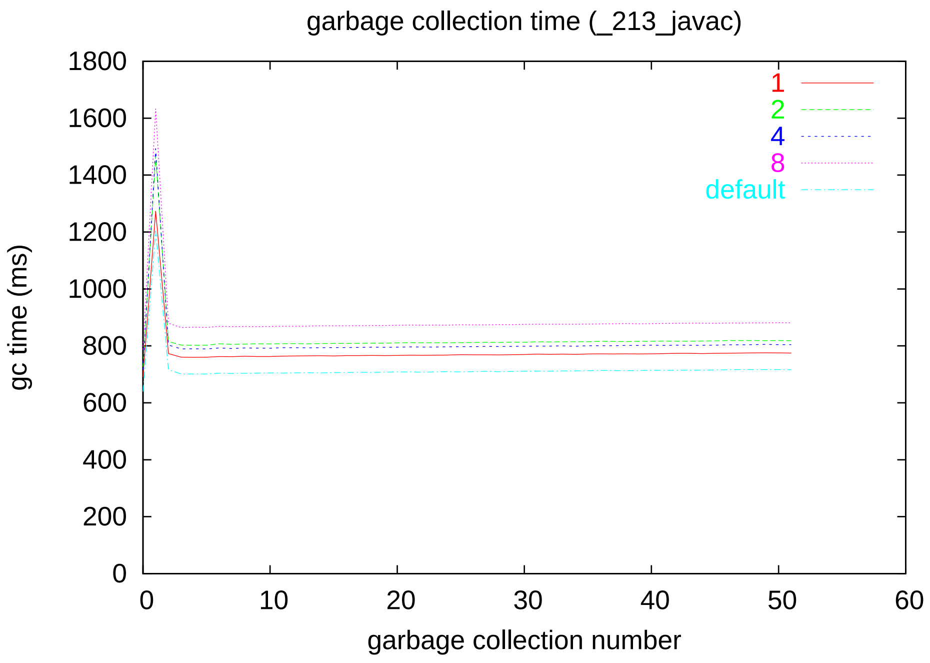


Figure 5.7: GC time when running *_213_javac*

code stubs, we used a compiler option to replace the allocated caller index by the *zero* index. When this option is enabled, calls do not execute the extra load instruction and profiling code stub, but still allocate CTB arrays for methods. For CFS and SpecJVM98 benchmarks, we found that usually the first run has some performance degradation when executing profiling code stubs (up to 9% except for _201_compress[7]), but the degradation is not significant upon reaching a stable state (between -2 to 3%). The performance of SpecJBB2000 is largely unaffected. Compared to not allocating CTB arrays at all (TIBs, however, are still expanded), the performance change is also very small.

benchmark	2	4	8
_201_compress	97.26%	99.99%	99.99%
_202_jess	0.93%	27.39%	41.10%
_209_db	97.39%	97.74%	99.99%
_213_javac	21.62%	64.25%	83.53%
_222_mpegaudio	40.81%	63.00%	78.38%
_227_mtrt	26.08%	73.82%	99.46%
_228_jack	48.51%	77.82%	86.01%

Table 5.12: Eliminated CTB loads by different inlining CTB sizes

Table 5.12 shows the percentages of eliminated CTB load instructions by different CTB inlining sizes. The experiment ignores call edges from and to RVM classes, and does not profile *static*, <*init*>, and interface methods. Each SpecJVM98 benchmark runs 10 times with input size 100 using the `SpecApplication` driver. The percentage of eliminated loads varies on different benchmarks. For example, loads of *_201_compress* and *_209_db* are mostly eliminated with an inlining size of 2, but _202_jess only has 41% eliminated even with an inlining size of 8. Other benchmarks have high elimination rates at inlining size 8. Eliminated loads did not cause significant performance changes. In our set of benchmarks, it seems that inlining more

[7]The first run of *_201_compress* does not promote enough methods to higher optimization levels.

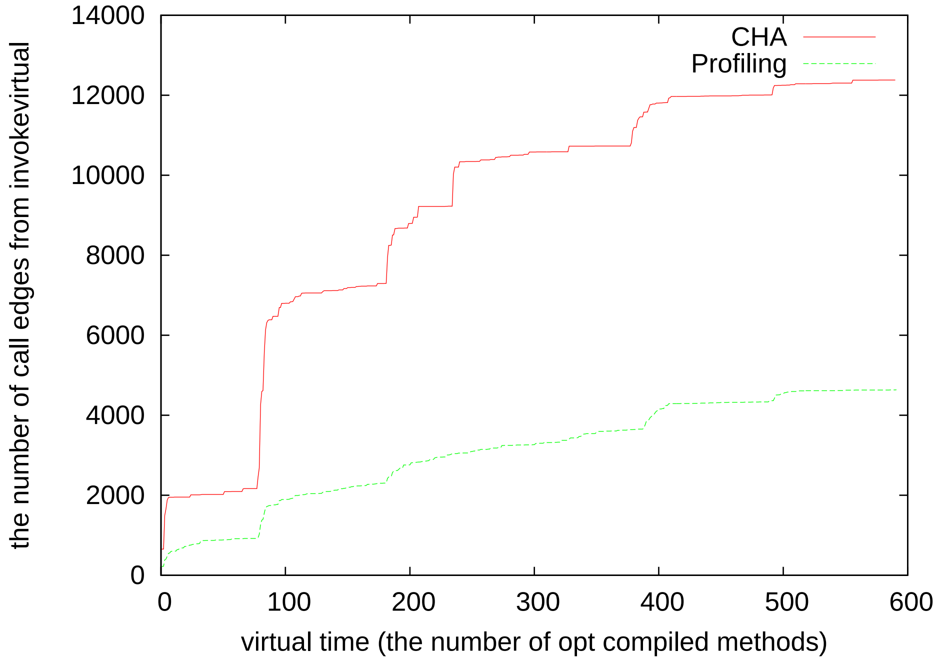


Figure 5.8: Growth of call graph at runtime (*_213_javac*)

CTB array elements does not result in further performance improvements.

Table 5.13 compares the sizes of profiled call graphs to these constructed by dynamic CHA. Each benchmark has two rows, the first row is the call graph size by CHA and the second row is the size by profiling. The third column (labelled as "ALL") gives the total number of call edges (application and library only). The number of call edges by CHA is same as Table 5.5. The row of "PROF" also has calculated percentages of call edges comparing to the row of "CHA". From the last column, we can see that profiled call graphs have 24% to 63% fewer *virtual* edges than CHA ones. The number of call edges from other call sites are similar because we used type analyses to compute them. Overall, profiling mechanism is able to reduce the total number of edges by 15% to 54% as shown in the third column. The reduction for the number of methods is not as significant as for the number of call edges.

benchmark		ALL		virtual	
_201_compress	CHA	1862		774	
	PROF	1446	78%	381	49%
_202_jess	CHA	4325		2070	
	PROF	3432	79%	1273	61%
_205_raytrace	CHA	3023		1751	
	PROF	2585	86%	1338	76%
_209_db	CHA	2177		941	
	PROF	1660	76%	506	54%
_213_javac	CHA	14936		12390	
	PROF	6917	46%	4645	37%
_222_mpegaudio	CHA	2520		1239	
	PROF	1940	77%	687	55%
_227_mtrt	CHA	3024		1751	
	PROF	2587	86%	1339	76%
_228_jack	CHA	4658		2278	
	PROF	3538	76%	1441	63%
SpecJBB2000	CHA	7186		4775	
	PROF	5517	77%	3251	68%
CFS	CHA	3385		2545	
	PROF	1776	52%	1018	40%

Table 5.13: Call graph comparison of CHA and profiling (application and library)

Call graph sizes shown before were collected at the end of benchmark runs. Consider applications of call graphs, it is more likely to be used at runtime by interprocedural analyses. Figure 5.8 shows the call graph size changes of *_213_javac* when the benchmark runs. The x-axis is the virtual time using the number of methods recompiled by the optimizing compiler. The y-axis is the number of call edges. The y-values at the end of x-axis is what reported in Table 5.13. From the figure, we can see that the sizes of call graphs constructed by different approaches have a similar ratio during the benchmark execution as the end of run. This confirms the consistent improvement of each call graph construction method.

A call graph client can use profiling mechanism with flexibility. For example, a client analysis could re-profile cold call edges to improve data-flow analysis results. After a client receives a call edge event, it performs propagation, then it can remove the call edge and require the VM to re-profile the same edge. If this call edge only executed once, future propagations will not pass data-flow information through the edge. This may improve the results of client analysis in the cost of more profiling overhead. However, type analyses cannot accomplish this task because call edge construction depends on class resolution, compilation or allocation events.

5.5 Related work

Static call graph construction for object-oriented programs focuses on approximating a set of types that a receiver of a polymorphic call site may have at runtime. Both CHA [DGC95] and RTA [BS96] are fast type analyses for method inlining and call graph construction. In a Java virtual machine, when type analyses are limited to initialized classes, we found that dynamic CHA leaves little room for improvement. Dynamic RTA is less effective than static one. The *instantiation-based type analysis* (ITA) is able to improve CHA call graphs by a small margin. However, three type analyses are not close to the limit as shown in Figure 5.13.

Reachability-based algorithms [Ste96,And94,SHR$^+$00,TP00] propagate types from allocation sites to receivers of polymorphic call sites along a program's control flow.

Assignments, method calls, field and array accesses may pass types from one variable to another. The analyses can either use a call graph built by CHA or RTA, then refine it, or build the call graph on the fly [RMR01].

Trampoline is a technique for generating a piece of self-modifying code on-the-fly. Java virtual machines heavily used this technique for implementing lazy compilation and class loading. Our call graph profiling stub is a self-modifying trampoline which pays the cost at the first-time execution.

In this section we have exhaustively studied call graph construction problem in Java virtual machines. We showed a general approach to deal with dynamic class loading and unresolved references in dynamic type analyses. A unique ITA is proposed for approximating call graphs of the bootimage. We have also proposed a profiling-based call graph construction mechanism, which builds most precise call graphs at runtime. Algorithms were implemented in Jikes RVM and evaluated using a set of Java benchmarks. An important characteristic of the dynamic call graph is that it supports speculative optimizations with invalidation backups.

Chapter 6
Online Interprocedural Type Analysis and Method Inlining

Using dynamic call graphs constructed by mechanisms in Chapter 5, we developped two reachability-based interprocedural type analysis, XTA and VTA, in Jikes RVM. The type analysis results are used for speculative method inlining.

This chapter also seeks to determine if more powerful dynamic type analyses could further improve inlining opportunities in a JIT compiler. To achieve this goal we developed a general *dynamic* type analysis framework which we have used for designing and implementing dynamic versions of several well-known static type analyses, including CHA, RTA, XTA and VTA.

Surprisingly, the simple dynamic CHA is nearly as good as an ideal type analysis for inlining **virtual method** calls. There is little room for further improvement. On the other hand, only a reachability-based interprocedural type analysis (VTA) is able to capture the majority of monomorphic **interface** calls.

We also found that memory overhead is the biggest issue with dynamic whole-program analyses. We used a generational garbage collector to reduce the impact of VTA data structures and measured performance improvement. We also present demand-driven approaches to reduce the memory overhead of dyanmic IPAs.

6.1 Introduction

Object-oriented programming languages encourage programmers to write small methods and compact classes so that the code is easy to read, modify, and maintain. Java programs exemplify this design idea: many tiny methods have only one line of code to access a field, return a hash value, or invoke another method. Design patterns [GHJV95] use class inheritance and virtual calls extensively to obtain great engineering benefits. Code instrumentation tools, such as AspectJ compilers [aspb, aspa], insert many small methods into instrumented programs. The downside of using small methods is that a program has to make frequent method calls. Object-oriented programs heavily rely on compilers to reduce calling overhead.

Efficient implementation of polymorphic calls has been studied extensively in the context of C++ [Dri01]. The Java programming language only allows single inheritance on normal classes, but allows multiple inheritance on interfaces. Virtual calls in Java can be categorized into two kinds: *virtual calls* on normal class types and *interface calls* on interfaces. Virtual calls can be implemented very efficiently by modern JIT compilers. Various techniques reduce the overhead of interface calls as well.

Even though the direct overhead of virtual calls is low, further performance improvement is often obtained from method inlining and optimizations on inlined code. Inlining creates larger code blocks for program analyses and improves the accuracy of intraprocedural analyses which must often handle method calls conservatively. Thus, method inlining is a very important part of a Java optimizer because it further reduces method call overhead and also increases other opportunities for optimizations.

A key step of method inlining is to decide which method(s) can be inlined at a call site. This can be achieved by using information conveyed via language constructs such as *final* and *private* declarations (which provide restrictions on which methods could be called), or the information can be gathered using a type analysis which determines which runtime types may be associated with a receiver, and hence which methods may be called. Another alternative is to profile targets of call sites. Inlining based on language constructs and type analyses results is conservative at analysis time and it supports direct inlining that maximizes optimization opportunities. In this chapter, we study method inlining using type analysis results.

Static type analyses for Java programs [DGC95, BS96, SHR+00, TP00] are not directly applicable to JIT compilers because of dynamic features of Java virtual machines. The type set of a variable might have new members as new classes are loaded and thus optimizations based on old results could be invalidated. Various techniques have been devised to use dynamic class hierarchy analysis for directly inlining in the presence of dynamic class loading and JIT compilation.

In this chapter we evaluate the effectiveness of several dynamic type analyses for method inlining in a Java virtual machine (Jikes RVM [AAB+00]). We built a common type analysis framework for expressing dynamic type analyses and used the results of these analyses for speculative inlining with invalidations. We then used this framework to perform a study of how many method calls can be inlined for the different varieties of type analyses.

We were also interested in finding the upper bound on how many calls that can be inlined, to determine if more accurate type analyses are required. To gather this information we used an efficient call graph profiling mechanism [QH04] to log call targets of each virtual call site. The logged information is used as an ideal type analysis for re-executing the benchmark. We compare the inlining results of other type analyses to the ideal one. In order to measure the maximum inlining potential of a type analysis, we also relaxed the size limit on inlining targets.

Our results were quite surprising. The simple CHA is nearly as good as the ideal type analysis for inlining virtual method calls and leaves little room for improvement. On the other hand, CHA is less effective for inlining interface calls. Further, we found that the majority of interface invocations are from a small number of hot call sites which are used in a very simple pattern.

In order to capture the monomorphic interface calls we developed *dynamic VTA*, which is a whole-program analysis. We analyzed the effectiveness and costs of this whole-program approach. We found that the main difficulty of such a dynamic whole-program analysis is that it requires large heap space for maintaining analysis data which must co-exist with application data in the heap. From our experience, we believe a demand-driven approach would make a dynamic interprocedural analysis practical in Java virtual machines and we suggest such an approach.

Our objective is to understand how well a dynamic type analysis can perform with respect to method inlining in a JIT compiler, and what opportunities there are for improvement. In this study, we made following contributions:

- A limit study of method inlining using dynamic type analyses on a set of standard Java benchmarks;
- Development and experience of an interprocedural reachability-based type analysis in a JIT environment;
- Interesting observations of speculative inlining and a proposal of demand-driven interprocedural type analyses.

Readers who are interested in the background of method inlining should read the Chapter 4. In this chapter, we describe the design of a common type analysis framework for speculative inlining in Section 6.2, The limit study results are also presented in this section. The whole-program VTA type analysis is described in Section 6.3 with experimental results. Related work is discussed in Section 6.4. Finally, in Section 6.5, we conclude with some observations and plans for future work.

6.2 A type analysis framework for method inlining

A *static* analysis is performed at compile-time and must make conservative assumptions that include all possible runtime executions. A static type analysis answers a basic question: what is the set of all possible runtime types of variable v at program point P. A *dynamic* type analysis is performed in a JIT environment, and therefore it is *time-sensitive*. It answers a query similar to a static one, except the answer is not for *all executions*, but for execution prior the time of answering the query. The results may change over program's execution. In order to use type analysis results for optimizations in a JIT environment, there are a few requirements we set for the analysis:

`dynamic:` it has to handle Java's dynamic features seamlessly, such as dynamic class loading, reference resolution, and JIT compilation;

`conservative:` analysis results must be correct at analysis time with respect to the executed part of the program;

`just-in-time:` the analysis should be able to notify clients when previous analysis results are about to change during execution.

A *dynamic* type analysis fits into a Java virtual machine without changing the lazy strategy of handling class loading and compilation. The *conservativeness* ensures optimizations based on analysis results are correct at the analysis time (it might be invalidated in the future). If the analysis can update its results *just-in-time*, it can be used for speculative optimizations with some invalidation mechanisms. Our objective is to design a type analysis framework supporting speculative inlining in a JIT compiler.

6.2.1 Framework structure

We designed a type analysis interface shown in Figure 6.1. In a Java method, a call site is uniquely identified by the method and a bytecode index. Given the method and bytecode index, the `getNodeId` method returns a node ID for further queries. The node ID allocation decides the granularity of different type analyses. For example, CHA and RTA use a single ID for all call sites, XTA allocates a node ID for all call sites in the same method, and VTA assigns different IDs to different call sites. The `lookupTargets` method returns an array of targets resolved by using reaching types of the node with a given callee method signature. The detailed lookup procedure is the same as virtual method lookup, defined by the JVM specification [LY96]. An inline oracle makes inline decisions according to the lookup results.

If the type analysis finds a monomorphic call site (with only one target), then the oracle decides to perform speculative inlining (using preexistence or code patching). It must register a dependency via the `checkAndRegisterDependency` method. A dependency says that, given a node and a callee method signature, a compiled method (*cm*) is valid only when the lookup results have one target that is the same as the parameter *target*.

After registering the dependency successfully, any change in the type set of the node causes verification of dependencies on this node. The `verifyDependency` method is called by the type analysis when the node has a new reaching type. For each dependency of the node, the verification procedure performs method lookup using the new reaching type and the callee method signature. If the lookup result is different from the target method of the dependency, the compiled method must be invalidated immediately.

```
public interface TypeAnalysis {
  public int getNodeId(VM_Method caller, int bcindex);
  public VM_Method[] lookupTargets(int nodeid, VM_Method callee);
  public boolean checkAndRegisterDependency(int nodeid,
                                            VM_Method callee,
                                            VM_CompiledMethod cm,
                                            VM_Method target);

  protected void verifyDependency(int nodeid, VM_Class newKls);
}
```

Figure 6.1: TypeAnalaysis interface

A *TypeAnalysis* implementation has to monitor system events such as class loading, method compilation, etc. We have implemented several type analyses as depicted in Figure 6.2[1]. CHA and RTA only differentiate classes that participate in the reaching type sets. We made a new variation of CHA, called ITA, to only allow classes with instances to participate in reaching types. XTA and VTA share many components. A special class, `IdealTypeAnalysis`, uses profiled results for the purpose of our limit study. All implementations satisfy the requirements defined at the beginning of this section. A client, `StaticInlineOracle`, uses the analysis results for

[1]JikesRVM has an implementation of dynamic CHA, we re-implemented it in our framework with little efforts.

speculative inlining.

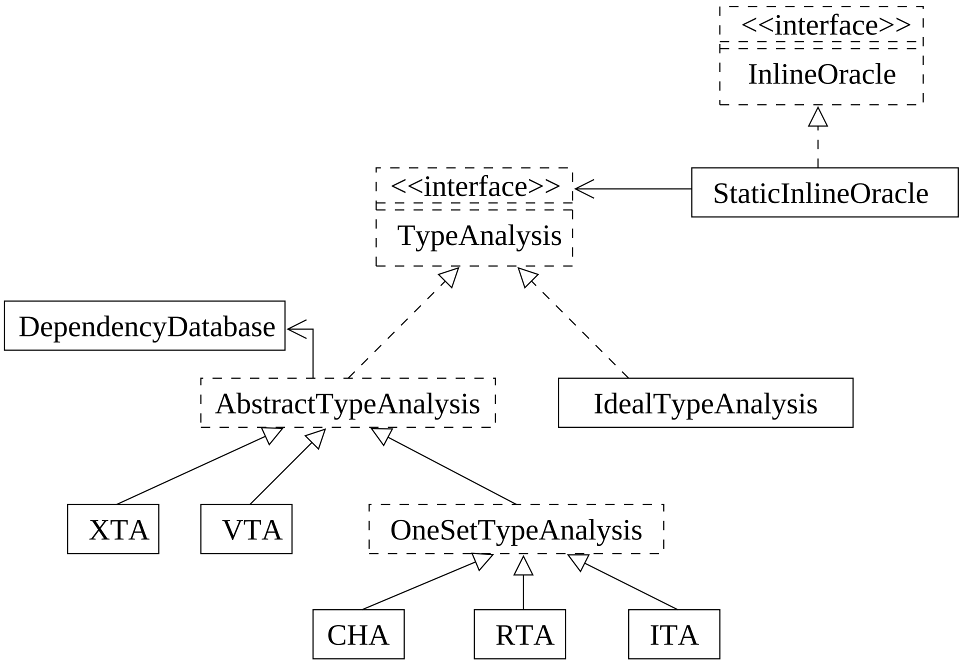


Figure 6.2: Type analysis framework diagram

6.2.2 A limit study of method inlining using dynamic type analyses

An ideal type analysis

To measure how precise a type analysis could be, we need an ideal type analysis for comparison. If a benchmark runs deterministically, we can profile targets in the first run, and then use the profiled targets as faked analysis results for the second run. We use an inexpensive call graph profiling mechanism [QH04] to gather call targets. An *IdealTypeAnalysis* parses the profiled targets for call sites, and the *lookupTargets* method returns profiled target(s) for a call site. The *IdealTypeAnalysis* uses CHA for call sites from the RVM boot image.

Experimental approach

An inline oracle has to balance the benefits and costs of inlining. Excessive inlining may blow up code size and slow down the execution. Therefore, a JIT compiler usually sets a size limit on inlined targets using some heuristics. Hazelwood and Grove [HG03] described the size heuristic used in Jikes RVM. For the purpose of our study, we would like to measure the maximum potential of a type analysis for method inlining without a size limit. However, inlining all call sites is not feasible. Instead, we only inline the most frequently executed call sites, without a size limit.

We implemented the framework and type analyses in JikesRVM v2.3.0. We used the *FastAdaptiveCopyMS* configuration for initial experiments since it is stable and can run all of our benchmarks. The configuration uses a copying mark-sweep collector.

Benchmarks

Our benchmark set includes the SpecJVM98 suite [speb], SpecJBB2000 [spea], a CFS subset evaluator from a data mining package Weka [wek], a simulator of certificate revocation schemes [cer], and a variation of the simulator interwoven with AspectJ code for detecting calls that return `null` on error conditions.

Table 6.1 summarizes dynamic characteristics of benchmark executions. We ignored call sites in the RVM code and Java libraries compiled into the boot image. Virtual and interface calls are measured separately. Columns labeled *total* report the total counts of invocations in each category. Columns labeled *#hottest* are numbers of hottest call sites, ranked in the top 100, whose invocations are more than 1% of *total* in Columns 2 and 5. Columns labeled *coverage* are percentages of invocations contributed by these hottest call sites.

It is interesting to point out that, for most of the benchmarks, the majority of invocations are from a small number of hot call sites. Less than 25 call sites exceed the 1% threshold. Only about half of the benchmarks have more than 1M interface invocations. These benchmarks have fewer than 10 hot interface call sites that contribute to more than 92% of invocations.

benchmark	invokevirtual			invokeinterface		
	total	#hottest	coverage	total	#hottest	coverage
_201_compress	2,191M	7	89%	0	N/A	N/A
_202_jess	964M	25	71%	0	N/A	N/A
_205_raytrace	2,837M	16	29%	0	N/A	N/A
_209_db	762M	8	99%	149M	5	99%
_213_javac	688M	10	20%	34M	5	92%
_222_mpegaudio	846M	25	80%	2M	11	98%
_228_jack	264M	22	74%	46M	11	93%
SpecJBB2000	8,162M	9	34%	146M	7	99%
CFS	639M	15	92%	0	N/A	N/A
simulator(orig)	44M	5	71%	0	N/A	N/A
simulator(aspects)	162M	13	72%	0	N/A	N/A

Table 6.1: Coverage of the hottest call sites

The `_213_javac` benchmark includes a large amount of auto-generated code. Invocation counts are spread over many call sites. `SpecJBB2000` has a large code base as well, and it runs much longer than other benchmarks. Hot call sites selected by our 1% threshold contribute only about 34% of total invocations.

A list of hottest call sites are provided to the inline oracle. The size limit is removed for call sites in the list. Thus, the inline oracle can exploit the potential of a type analysis as much as possible.

As we discussed in Chapter 4, a virtual call site can be inlined using different techniques:

- *direct*: direct inlining if the callee method is *private* or *final*;
- *preex*: direct inlining with invalidation checks if the receiver can be proved to be preexistent prior method calls;
- *cp*: guarded inlining with code patching;

- *mt* or *ct*: guarded inlining with method or class tests.

If a call site is currently monomorphic according to the analysis results, guards are chosen as a command line option. It can be code patching or method/class tests. For our experiment we used code patching since it has less runtime overhead.

Monomorphic interface calls can be directly inlined if the receiver is preexistent, or inlined with guards. We found that, in our benchmark set, receivers of nearly all hot interface calls cannot be proved to be preexistent by an invariant argument analysis. In our results, we omit the *preex* category for interface calls. We also performed another experiment where the inline oracle inlined polymorphic call sites (guarded by method or class tests) that had 1 or 2 targets resolved using type analysis results. However, this did not lead to significantly more inlined calls (only `_213_javac` has a 2% increase). Thus, we do not inline polymorphic calls in our experiment reported here.

Limit study results

Table 6.2 compares the results of dynamic *CHA* and *IdealTypeAnalysis*. Each benchmark has two rows: `ideal` and `cha`, showing dynamic counts of inlined calls using different type analyses. Virtual and interface calls are presented separately. Column *total* is the count of invocations in each category. In the *virtual* category, dynamic CHA did nearly as perfect a job as the ideal type analysis in most benchmarks, except `_213_javac` and `simulator(aspects)`. On these benchmarks, the majority of dynamic invocations are contributed by monomorphic call sites. The sum of *direct*, *preex* and *cp* is close to the coverage in Table 6.1. `_213_javac` leaves a small gap between *cha* and *ideal*. In the *interface* category, column 8 shows that a large portion of interface invocations are from monomorphic call sites as well. Dynamic CHA is ineffective on inlining interface calls. Furthermore, the other two simple type analyses, RTA and ITA, did not improve the results of inlining interface calls because common interfaces are implemented by different classes that are likely to be instantiated.

		virtual				interface		
		total	direct	preex	cp	total	cp	mt
_201_compress	ideal	2,191M	99%	0	0	0	0	0
	cha	2,191M	99%	0	0	0	0	0
_202_jess	ideal	994M	58%	6%	21%	7M	0	0
	cha	994M	58%	6%	21%	7M	0	0
_205_raytrace	ideal	2,837M	0	50%	41%	0	0	0
	cha	2,837M	0	50%	41%	0	0	0
_209_db	ideal	762M	31%	0	67%	150M	99%	0
	cha	762M	31%	0	67%	150M	0	0
_213_javac	ideal	701M	28%	7%	15%	35M	95%	0
	cha	701M	27%	7%	10%	35M	0	0
_222_mpegaudio	ideal	846M	73%	3%	0	2M	57%	0
	cha	846M	73%	3%	0	2M	57%	0
_228_jack	ideal	258M	13%	16%	39%	46M	86%	0
	cha	258M	12%	15%	39%	46M	25%	8%
SpecJBB2000	ideal	8,250M	32%	29%	11%	148M	99%	0
	cha	8,119M	32%	29%	12%	146M	0	0
CFS	ideal	639M	38%	6%	52%	0	0	0
	cha	639M	38%	6%	52%	0	0	0
simulator	ideal	44M	99%	0	0	0	0	0
(original)	cha	44M	99%	0	0	0	0	0
simulator	ideal	162M	11%	19%	53%	0	0	0
(aspects)	cha	162M	11%	0	53%	0	0	0

Table 6.2: Limit study of method inlining using type analyses

Discussion

`Simulator(aspects)` is an interesting benchmark. Injecting AspectJ advice code increases the number of invocations and changes inlining behaviors dramatically. In the original benchmark, nearly all virtual calls are monomorphic and can be directly inlined. With aspects, dynamic CHA misses all monomorphic calls in the *preex* category. After looking at the benchmark closely, we found this is due to the generic implementation of pointcuts.

The pointcut implementation boxes primitive values in objects and passes them to AspectJ libraries. The value is then unboxed after the library call. The original code for unboxing looks like

```
int intValue(Object v) {
  if (v instanceof Number)
    return ((Number)v).intValue();
  ......
}
```

The single call site of `((Number)v).intValue()` contributes 19% *preex* invocations. Dynamic CHA failed to inline this call site because the `Number` class has several subclasses, `Integer`, `Double`, and `Long`, and the call site is identified as polymorphic.

This particular problem can be solved in two ways: 1) use a context-sensitive reachability-based type analysis, or 2) change the implementation of unboxing to facilitate the type analysis. We changed the method to use a tighter type, `Integer`, in the type cast expression, then the call site becomes directly inlineable.

Since the number of hot interface call sites is small, we investigated them one by one. It turns out these hot interface calls are used in a similar pattern:

```
// <TYPE> is java.util.Vector, java.util.Hashtable, etc.
Enumeration e = <TYPE>.elements();
......
while (e.hasMoreElements())
  index[i++] = (Entry)e.nextElement();
```

`Enumeration` is an interface in `java.util` package. The *while* loop makes two or more interface calls for enumerating elements of underlying data structures. Dynamic CHA assumes all implementations of the interface are in the runtime type set of `e`, although each `<TYPE>` class returns a specific implementation. Without interprocedural information or inlining the `<TYPE>.elements()` method, a type analysis cannot produce precise type information of `e`. Therefore, these interface call sites cannot be inlined by using dynamic CHA.

From this limit study, we conclude that:

- most virtual calls in standard Java benchmarks are monomorphic;
- dynamic CHA is nearly perfect for inlining virtual calls;
- dynamic CHA is ineffective on inlining interface calls;
- to assist compiler optimizations, a programmer should use precise types when it does not sacrifice other engineering benefits;
- a large percentage of interface calls are monomorphic and used in a simple pattern, but it requires an interprocedural analysis to discover the precise type of the receiver.

6.3 Dynamic reachability-based type analysis

In Section 6.2, we presented a type analysis framework for supporting speculative inlining. We also presented the results of our limit study of method inlining which showed that dynamic CHA is not strong enough for inlining interface calls. In this section, we present an interprocedural, reachability-based, type analysis that is suitable for inlining interface calls.

There are two different approaches to performing a dynamic interprocedural analysis in a Java virtual machine. A whole-program analysis analyzes all classes and methods that can participate the program execution. A demand-driven analysis only analyzes the part of code related to a request. We start with the whole-program approach.

We designed and implemented a dynamic version XTA in our previous work [QH04] as an example of how to deal with dynamic class loading and reference resolution. However, due to lack of intraprocedural data-flow information, XTA results are very coarse. Although the computed type sets are smaller than ones from CHA, it still could not recognize important monomorphic interface call sites. From the method inlining study, we found XTA results were no better than dynamic CHA.

Design

VTA [SHR$^+$00] uses intraprocedural data flow information to propagate type sets. Given a Java program (all application and library classes), static VTA constructs a directed type flow graph $G = (V, E, \tau)$ where:

- V is a set of nodes, representing local variables, method formals and returns, static and instance fields, and array elements;
- E is a set of directed edges between nodes, an edge $a \rightarrow b$ represents an assignment of a's value to b;
- $\tau : V \rightarrow T$ is a map from a node to a set of types (classes).

Static VTA has two phases. On phase 1, a constrain collector performs one-pass scan of the program and constructs a VTA graph. Phase 2 propagates types $\tau(V)$ to all reachable nodes in the graph.

Dynamic VTA can also take advantage of rich runtime type information. A dynamic VTA node has a declaring type inferenced by JikesRVM's optimizaing compiler. When propagating types through a node v, only when v's declaring type is resolved and its subtypes can become part of $\tau(v)$.

We use the same approach outlined in [QH04] to adopt the static VTA to a JIT compiler. In the whole-program approach, the constraint collector monitors method compilation events at runtime. Before a method is compiled, the constraint collector parses the bytecode and creates VTA edges. The collector uses the front-end of the optimizing compiler in Jikes RVM, which converts bytecode to a three address intermediate representation, HIR. Several optimizations are performed during translation. An HIR operand has a declaring type.

Dynamic VTA analysis is driven by events from JIT compilers and class loaders. Figure 6.3 shows the flow of events. In the dotted box are the three modules of dynamic VTA analysis: VTA graphs, the analysis (include constraint collector), and dependency databases. The JIT compilers notify the analysis by channel 1 that a method is about to be compiled. The analysis scans the bytecode of the method body and, for each *new* instruction with a resolved type, the analysis adds the type into the reachable type set of the method via channel 3; otherwise it registers a dependency on the unresolved type reference for the method via channel 4. Similarly for field accesses, if the field reference can be resolved without triggering class loading, the analysis adds a directed edge into the graph via channel 3; otherwise, it registers a dependency on unresolved field reference for the method. A call graph constructor 5 could add new edges to the graph by channel 2. Whenever a type reference or field reference gets resolved, the dependency databases are notified (by channel 5), and registered dependencies on resolved references are resolved to new reachable types or new edges of the graph.

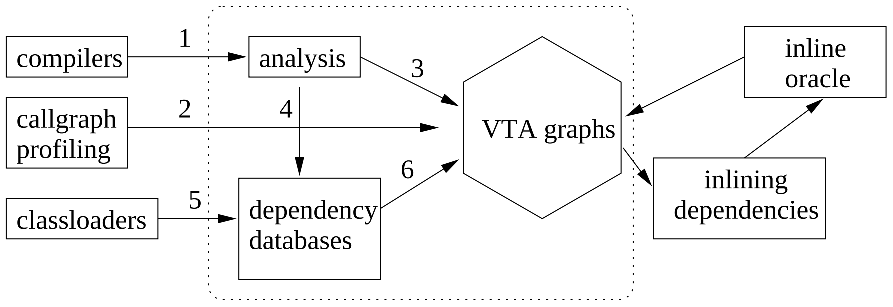


Figure 6.3: Model of VTA events

Many system events can change the VTA graph. Whenever the graph is changed (either the graph has an new edge, or a node has a new reaching type), a propagator propagates type sets of nodes (related to changes) until no further change occurs. Whenever the reaching type set of a node has a new member, the analysis verifies dependencies on this node registered by inlining oracles (see Section 6.2). The oracle has a chance to perform invalidation if inlining assumptions are violated.

Propagations

To support speculative optimizations, the analysis must keep the results up-to-date whenever a client makes queries. An *eager* approach propagates new types whenever the VTA graph is changed. The second approach is to cache graph changes when collecting constraints of a method, and *batch* propagations at the end of constraint collection. The third approach, as suggested in [HDH04, PS01], propagates types only when a client makes queries on nodes. However, the analysis needs to keep a list of nodes which type sets are used for speculative optimizations. Whenever the VTA graph changes, the analysis has to perform demand-driven propagation on listed nodes to verify assumptions are not invalidated. In our study, we found both eager and batch propagations are efficient, with respect to the total execution time of each benchmark. Javac takes up to 1.7 seconds and other benchmarks take less than 1 second.

Implementation

The main data structure of VTA is a type flow graph. Previous work [LH03] showed that the ordinary set and map implementations from JDK are not scalable. We used a similar implementation of integer sets as in [LH03]. In addition, we have implemented a special hash table using primitive integers as keys.

Effectiveness of dynamic VTA

Not surprisingly, VTA is able to handle the simple pattern of interface calls in our benchmarks set. Table 6.3 compares dynamic counts of inlined interface calls. We omitted benchmarks with few interface calls. Dynamic VTA is able to catch all monomorphic interface calls and allows them to be inlined by using code patching.

benchmark	Ideal(cp)	VTA(cp)
_209_db	99%	99%
_213_javac	95%	95%
_228_jack	86%	86%
SpecJBB2000	99%	99%

Table 6.3: Comparison of VTA and IdealTypeAnalysis for inlining interface calls

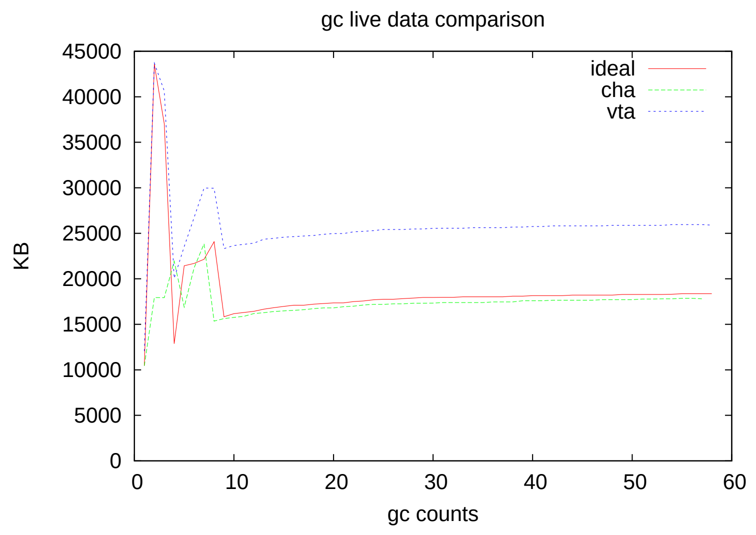
Memory overhead of whole-program VTA

Although dynamic VTA allows the JIT compiler to utilize maximum inlining opportunities, the cost of whole-program VTA is also high. Table 6.4 is a rough approximation of memory footprint of VTA graphs for four benchmarks. A VTA graph has three big pieces: a node set, an edge set, and type sets. Columns 2 to 4 display numbers of sets and memory footprints. The last column is the total memory footprint of three pieces.

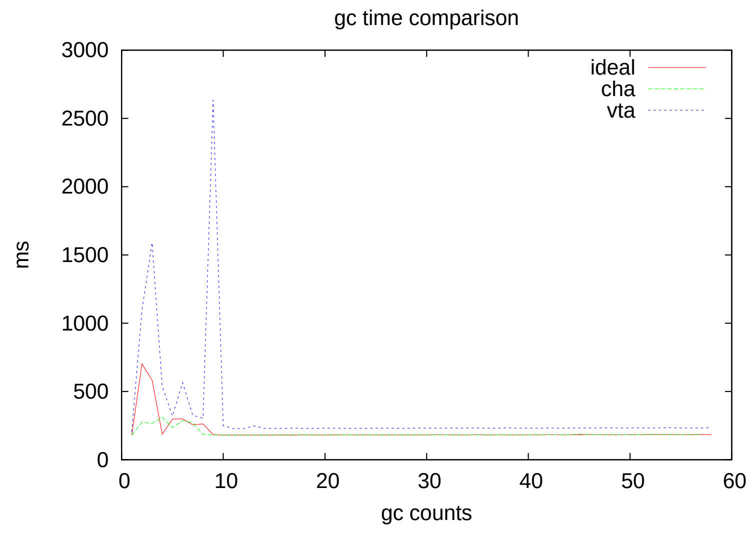
benchmark	nodes	(size)	edges	(size)	typesets	total
_209_db	4,326	(0.13M)	5,419	(0.53M)	0.31M	0.97M
_213_javac	11,933	(0.37M)	32,992	(2.54M)	1.28M	4.19M
_228_jack	7,261	(0.22M)	10,403	(1.00M)	0.64M	1.86M
SpecJBB2000	11,119	(0.34M)	14,831	(1.54M)	0.80M	2.68M

Table 6.4: VTA graph sizes

Figure 6.4(a) compares sizes of live data for *IdealTypeAnalysis*, *CHA*, and *VTA*, using `_213_javac` as the example. Heap occupation has been increased by one-third (about 10M) for VTA graphs. Also VTA data are lived through application execution. Figure 6.4(b) depicts time spent on each GC using a copying collector. It is clear that the whole-program interprocedural analysis has a very high memory overhead.



(a) live data after each GC



(b) time spent on each GC

Figure 6.4: GC behaviors affected by VTA (*_213_javac*)

Performance comparison

Table 6.5 shows our preliminary performance measurement using a *copying mark-sweep* collector. Data were collected on a laptop with a Pentium M 1.5G processor and 1G memory, running Linux kernel 2.6.10. We took the best run of 10 runs of SpecJVM98 benchmarks. SpecJBB2000 is measured by its throughput score (higher is better). VTA has some speedups on *_209_db* and *_228_jack*. `_213_javac` and `SpecJBB2000` slowed down due to heavier GC workload introduced by VTA graphs,

benchmark	CHA	VTA	speedups
_209_db	10.704s	10.448s	2.5%
_213_javac	4.006s	4.048s	-1.0%
_228_jack	2.638s	2.635s	0.1%
SpecJBB2000	12093	9693	-24.8%

Table 6.5: Performance comparison using a copying mark-sweep collector

Recently we switched to a *generational mark-sweep* collector, which promotes most of VTA graph objects to old generations. The impact of GC has been reduced. Table 6.6 compares the best run of 10 runs of two benchmarks. Unfortunately, both `_209_db` and `SpecJBB2000` trigger bugs in the *generational mark-sweep* collector in the version of JikesRVM that we are using for our implementation.[2]

benchmark	CHA	VTA	speedups
_213_javac	3.924s	3.900s	0.6%
_228_jack	2.514s	2.460s	2.1%

Table 6.6: Performance comparison using a generational mark-sweep collector

[2]It is possible that this issue will be resolved when we upgrade our implementation to the latest version of JikesRVM.

6.4 Related work

We discussed some related work of method inlining in Chapter 4. The section discusses additional related work on the topic.

Ishizaki et al. [IKY$^+$00] conducted an extensive study of dynamic devirtualization techniques for Java programs. In their experiments, size limits were put on inlined targets, and techniques using dynamic CHA were shown to inline about 46% of virtual calls (execution counts). Our study answers the question of what is the limit of method inlining using different type analyses. By lifting the size limit on hottest call sites, we were able to understand the maximum inlining potential using a type analysis. Our limit study shows that CHA is nearly as perfect as an ideal type analysis.

Pechtchanski and Sarkar [PS01] presented a framework for dynamic optimistic interprocedural analysis (DOIT) in a JIT environment. For each method, the DOIT analysis builds a value graph similar to a VTA graph. However, due to lack of a complete dynamic call graph, DOIT does not track type flow between method calls (parameters and returns). Instead, it uses conservative subtypes of declaring types of method parameters and returns. DOIT is good at obtaining precise type information for fields whose values are assigned in one method and used by another method. Our work focused on limit study of method inlining using type analyses, including online interprocedural analyses based on dynamic call graphs. Our results independently confirms that dynamic CHA is effective for inlining virtual calls in Java programs.

Profiled-directed inlining is effective to identify profitable inlining targets at polymorphic call sites. However, profile-directed inlining requires runtime tests to guard the inlining target. Our focus is on exploiting unguarded inlining opportunities exposed by type analyses.

6.5 Discussion

In this chapter we have presented a study on the limits of speculative inlining. Somewhat to our surprise we found that using dynamic CHA for speculative inlining is almost as good as using an "ideal" analysis, for inlining virtual method calls. However, for even simple uses of interface calls, none of dynamic CHA, RTA, XTA or ITA gives enough information for determining that interface calls are monomorphic. Rather, to detect these opportunities, one requires a stronger type analysis and we presented a dynamic version of VTA for this purpose.

Our experiments with our dynamic VTA do show that it provides detailed enough type information to identify inlining opportunities for interface calls in our benchmark set. However, we also note that the memory overhead of our whole program approach to dynamic VTA is quite large, and we suggest an alternative demand-driven approach.

In addition to these main contributions of this chapter, we also made several other general observations about speculative method inlining.

Observation 1

The conventional wisdom is that inlining increases optimization opportunities. However, in the presence of speculative optimizations, inlining may reduce optimization opportunities as well. Figure 6.5 shows such an example. In Figure 6.5(a), the method `Foo.m()` is declared as *virtual*, but not overridden. Thus the call site in the `child` method is a candidate of direct inlining based on the receiver's preexistence prior method call (Figure 6.5(b)). However, if a compiler inlines `child()` into `parent()`, and the receiver of `foo.m()` is not preexistent prior the `parent()`, the call site can only be inlined with a guard as in Figure 6.5(c). Since the frequency of calling `foo.m()` is much more than calling `child()`, the performance of `parent()` might not be maximized. This pattern did happen in the *_213_javac* benchmark.

The above contradiction could be resolved by using on-stack replacement technology [FQ03, Urs92] or thin guards [AR02]. Indeed, method invalidation performs on-stack replacement at method entries. A compiler can insert a general on-stack replacement point after the statement *Foo f = getfiled* with a condition that *Foo.m* is currently final. The compiler can directly inline the body of *Foo.m* into the loop.

```
parent() {
  Foo f = getfield
  this.child(f);
}

child(Foo foo) {
  while(cond)
    foo.m()
}
```

(a) source

```
parent() {
 Foo f = getfield
 this.child(f);
}

child(Foo foo) {
 while (cond)
  inlined Foo.m(foo)
}
```

(b) inline 'Foo.m' only

```
parent(){
 Foo f = getfield
 while(cond)
  if (Foo.m is currently final)
   inlined Foo.m(f)
  else
   Foo.m(f)
}
```

(c) inline 'child', then 'Foo.m'

Figure 6.5: An example where inlining can reduce optimization opportunities

Observation 2

Our second observation is that inlining decisions may be affected by library implementations. A Java virtual machine is bundled with a specific implementation of Java class library. For example, the GNU classpath [cla] is an open-source implementation of Java libraries and used by many open source Java virtual machines, including Jikes RVM. The implementation of `Hashtable.elements()` in the GNU classpath (version 0.07) returns objects of a single type `Hashtable$Enumerator`. The implementation in Sun's JDK 1.4.2_04, however, may return objects of `Hashtable$EmptyEnumerator` and `Hashtable$Enumerator`. Several hot interface call sites in our benchmark set would not be directly inlined if using Sun's JDK.

Demand-driven propagation

VTA graphs of large benchmarks (e.g., *_213_javac* and *SpecJBB2000*) have large number of nodes. However, there are only a few hundreds nodes whose type sets are used for inlining. Both batch and eager propagations save type sets for all nodes. The demand-driven propagation [PS01] is attractive since it does not require the analysis to save type sets for intermediate nodes. The analysis can do a depth-first search on the VTA graph to find all reachable types of a node. However, in order to support speculative inlining, the analysis has to perform depth-first search on all receiver nodes of speculatively inlined sites, whenever the VTA graph changes. We are planning to implement this approach and measure the performance.

Outline of a demand-driven analysis

The whole-program interprocedural analysis analyzes all compiled methods. Although the analysis itself is reasonably fast, the VTA graph representation takes a lot of memory space. Because Jikes RVM shares the same heap with applications, more live data triggers more frequent garbage collections and increases GC time in each collection. Section 6.2 shows that only a few call sites are hot in each benchmark. If the analysis only analyzes code related to these hot call sites, the time and space overhead could be reduced.

Demand-driven analysis [VR01,HT01] itself is an interesting research topic. It has been explored in the context of static points-to analysis. It is attractive to perform incremental, demand-driven IPA at runtime in a JIT environment.

We are developing a demand-driven VTA type analysis. We outline the requirement and design of the demand-driven analysis to support speculative inlining of interface calls. Given a receiver variable `e`, the compiler would decide if it is interesting to perform the type analysis on `e`. A simple heuristic is that `e`'s declaring type is a common interface, such as *Enumeration* and *Iterator*, and there are interface invocations in loops made on the variable. Once the compiler decides to analyze `e`, the demand-driven analysis would look at the source of `e`. If there is a single definition of `e` in the method, it continues looking at where the value of e comes from. There are several possible cases, we describe each case and solution:

- `e` is passed in as a parameter. In this case, the analysis gives up since it is too expensive to analyze all callers of the method;
- `e` gets its value from a field or an array element. It requires a whole-program analysis in order to know all writes to the field or array element. Field analysis [GRS00] could be used if it is available, otherwise, the analysis gives up in this case;

- `e`'s value is returned from a method call. This is the case the analysis tries to analyze further, and it performs the following steps:
  - is there only one target of the call site? if so, then
  - can the type analysis get a precise type of return objects in the target method?

  If the analysis successfully finds the precise type of `e`, it then registers a dependency that assumes the analyzed method is the only target of the call site.

The above approach could successfully inline the common interface call patterns we found in our benchmarks. However, it is fairly restricted to these patterns. We believe the right approach should look at more patterns used in the real applications and analyze important ones case by case.

Future work

Based on this study we have concluded that a type analysis for *invokeinterface*s is an important area of research, and we are currently working on a demand-driven analysis and compact graph representation to reduce the costs of dynamic VTA. We are also looking at more applications of dynamic interprocedural analysis in JIT compilers. A new research topic is to investigate the effectiveness of compiler optimizations on different design patterns.

Chapter 7
Conclusions and Future Work

In this chapter, we summarize the thesis in Section 7.1. In Section 7.2, we discuss several potential research directions from the thesis work.

7.1 Conclusions

In this thesis we have explored runtime instrumentation and dynamic program analysis in a Java virtual machine. The objective is to investigate the opportunities and challenges of program analysis and optimizations in a runtime system for languages such as Java and C#.

In Chapter 3, we presented a novel dynamic memory allocation scheme in garbage collectors. The new scheme uses write barriers to gain the benefits of stack allocation, but avoiding the requirement of an escape analysis that is not largely available in JIT environments. In our allocator, an allocation site was dynamically attributed as escaping and non-escaping based on its execution history. In an escaping analysis, such a property is derived from programs and langauge constraints, and is often over-conservative.

We presented the overall idea, and provided details of a prototype design. In our implementation, we carefully measured the benchmark behaviors and optimized the instrumentation. Our quantative meansurement gave encouraging results. In the best case of *_227_mtrt*, the number of collections was reduced from 7 to 1.

A JIT environment allows speculative optimizations based on runtime invalidation techniques. In Chapter 4, we reviewed existing state-of-the-art work in this field. It covers simple runtime checks to complicated on-stack-replacement. We also presented an improvement and implementation of a new on-stack replacement mechanism in a real Java virtual machine. All these techniques are necessary to utilize more advanced program analyses for speculative optimizations.

A necessary step of interprocedural analyses is to construct call graphs efficiently. In Chapter 5, we did a thorough study of dynamic call graph construction problem in Java virtual machines. We showed a general approach of handling dynamic class loading and unresolved references in a dynamic program analysis. A call graph profiler, using trampoline, builds most precise call graphs with small runtime overhead. Type analysis and profiler can be coupled to make trade-off between efficiency and precision. We implemented and evaluated algorithms in Jikes RVM on a set of standard benchmarks.

It is relatively easy to adapt a static intraprocedural analysis to a JIT compiler since methods are compilation units in both envoriments. Interprocedural analyses, however, are largely unexplored in JIT compilers.

In Chapter 6, we presented the design and implementation of several dynamic interprocedural type analyses. We used method inlining as an experiment to study how to use analysis results for speculative optimizations. We found that a dynamic analysis only needs to focus on a narrow region of code, and it has quite different effects comparing to its static counterpart.

We examined necessary techniques to create maximum inlining opportunities. A dynamic interprocedural type analysis has been developed in Jikes RVM. We studied the costs and gains of a whole program analysis approach, and pointed out issues of the approach. In this study, we also made some interesting observations about dynamic program analysis and speculative inlining.

7.2 Future work

In this section, we pointed out some future research directions following the thesis work.

7.2.1 Dynamic interprocedural analyses and speculative optimizations

Dynamic interprocedural analysis is still not a well-understood research field. The whole program approach used in static analysis is unlikely to succeed because the cost of maintain intermediate data structures is too high in a Java virtual machine. The 90/10 rule applies on Java programs as well. A successful dynamic analysis should spend limited resource on the small percentage of hot code regions. On-demand approach might succeed in a JIT environment.

We point out some possible dynamic interprocedural analyses, but not exhaustive: interprocedural type analysis results can be used to inline interface call sites, and eliminate unnecessary type checks; escape analysis enables new garbage collection schemes, such as regions, connectivity-based GC, etc.; value range analysis allows bounds check elimination.

Advanced interprocedural analysis would create new optimization opportunities. We need to explore new kinds of speculative optimizations in a JIT environment. The current OSR prototype is a heavyweight implementation. It requires fine-grained tuning to alleviate the cost of backup mechanism.

7.2.2 Online escape analysis

Ruf [Ruf00] proposed a static escape analysis for synchronization removal in the Marmot compiler [FKR+98]. A similar analysis [Ste00] could be used for creating thread-local heaps. Not like other escape analyses [CGS+99, Bla99, WR99], Ruf's analysis does not assume that an object is escaping because it is reachable from a static field. Only objects that may be accessed by multiple threads are considered escaping. Ruf's analysis relies on a conservative call graph and annotates methods that may be accessed by multiple threads.

We could develop an online escape analysis similar to the one suggested in [Ste00] using our dynamic call graph constructors. There are several ways to improve the efficiency and precision of the dynamic analysis:

1. track thread creation and annotate call graph with live threads;

2. since the runtime system knows exact lifetime of a thread, an object accessed by threads with non-overlapping lifetime can still be considered as thread-local;

Using the hybrid of dynamic analysis and instrumentation, it is possible to develop an efficient and effective online escape analysis. However, applications of analysis results may require new invalidation techniques to ensure correct execution when old results were invalidated.

7.2.3 Improve the efficiency of on-stack-replacement

Our implementation of on-stack-replacement (Section 4.4) requires an OsrPoint instruction that uses all variables in the source program. It may limit the precision of intraprocedural analyses, such as constant propagation, common subexpression elimination, etc. One future research direction is to study how to reduce side-effects of on-stack-replacement on regular program optimizations.

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Appendix A
Dynamic metrics of benchmarks

Table A.1 lists six important metrics of benchmarks used in this thesis. The data come from the webpage of Sable research group at McGill University[1]. We use short names for benchmarks due to the limit of page width. Each row compares one metric of all benchmarks. We also use different metric names from the original source:

loaded.classes is the number of loaded application classes (a.k.a *size.appLoadedClasses.value*);

size.load is the number of bytecode instructions of loaded application classes (a.k.a *size.appLoad.value*);

size.exec is the number of bytecode instructions that are executed at least once (a.k.a *size.appRun.value*);

poly.sites is the number of different virtual call sites executed in application code (a.k.a *polymorphism.appCallSites.value*);

poly.polyrate is the percentage of call sites that have multiple targets (a.k.a *polymorphism.appTargetPolyDensity.value*);

mem.alcrate is the allocation rate as the number of bytes allocated per killo bytecode executed (a.k.a *memory.byteAppAllocationDensity.value*);

[1]http://www.sable.mcgill.ca/metrics

metric	comp	jess	db	javac	mpeg	mtrt	jack	soot
loaded.classes	22	158	14	175	62	35	66	522
size.load	6,555	22,370	6,436	44,664	38,484	11,193	23,424	45,278
size.exec	5,084	11,634	4,546	26,267	34,975	9,460	18,721	23,850
poly.sites	54	737	128	2617	326	939	1124	2877
poly.polyrate	0.019	0.011	0	0.103	0.037	0.005	0.010	0.049
mem.alcrate	11.096	294.895	24.824	131.824	0.431	91.766	313.793	167.476

Table A.1: Selected dynamic metrics of benchmarks