Points-to Analysis for Java Using Annotated Constraints

Atanas Rountev Ana Milanova Barbara G. Ryder

Department of Computer Science
Rutgers University
New Brunswick, NJ 08901
{rountev,milanova,ryder}@cs.rutgers.edu

ABSTRACT

The goal of points-to analysis for Java is to determine the set of objects pointed to by a reference variable or a reference object field. This information has a wide variety of client applications in optimizing compilers and software engineering tools. In this paper we present a points-to analysis for Java based on Andersen's points-to analysis for C [5]. We implement the analysis by using a constraint-based approach which employs *annotated inclusion constraints*. Constraint annotations allow us to model precisely and efficiently the semantics of virtual calls and the flow of values through object fields. By solving systems of annotated inclusion constraints, we have been able to perform practical and precise points-to analysis for Java.

We evaluate the performance of the analysis on a large set of Java programs. Our experiments show that the analysis runs in practical time and space. We also show that the points-to solution has significant impact on clients such as object read-write information, call graph construction, virtual call resolution, synchronization removal, and stack-based object allocation. Our results demonstrate that the analysis is a realistic candidate for a relatively precise, practical, general-purpose points-to analysis for Java.

1. INTRODUCTION

Performance improvement through the use of compiler technology is important for making Java a viable choice for production-strength software. In addition, the development of large Java software systems requires strong support from software engineering tools for program understanding, maintenance, and testing. Both optimizing compilers and software engineering tools employ various static analyses to determine properties of run-time program behavior. One fundamental static analysis is *points-to analysis*. For Java, points-to analysis determines the set of objects whose addresses may be stored in a given reference variable or reference object field. By computing such *points-to sets* for variables and fields, the analysis constructs an abstraction of the run-time memory states of the analyzed program. This abstraction is typically represented by one or more *points-to graphs*. (An example of a points-to graph is shown in Figure 1, which is discussed later.)

Points-to analysis enables a variety of other analyses—for example, *side-effect analysis*, which determines the memory locations that may be modified by the execution of a statement, and *def-use analysis*, which identifies pairs of statements that set the value of a memory location and subsequently use that value. Such analyses are needed by compilers to perform well-known optimizations such as code motion and partial redundancy elimination. These analyses are also important in the context of software engineering tools: for example, def-use analysis is needed for program slicing and data-flow-based testing. Points-to analysis is a crucial prerequisite for employing these analyses and optimizations.

In addition to enabling other analyses, points-to analysis can be used directly in optimizing Java compilers to perform a variety of popular optimizations such as virtual call resolution, removal of unnecessary synchronization, and stack-based object allocation. Typically, each of these optimizations is based on a *specialized* analysis designed for the purpose of this specific optimization. Thus, compilers that employ multiple optimizations need to implement many different analyses. In contrast, using a single *general-purpose* points-to analysis can enable several different optimizations. Furthermore, the cost of the analysis can be amortized across many client optimizations, and the development effort to implement the optimizations can be significantly reduced.

Because of the many applications of points-to analysis, it is important to investigate approaches for precise and efficient computation of points-to information. In this paper we define and evaluate a points-to analysis for Java which is based on Andersen's points-to analysis for C [5], with all extensions necessary to handle object-oriented features.

Andersen's analysis is a relatively precise flow- and context-insensitive analysis[1] with cubic worst-case complexity. Despite this complexity, previous work has shown that certain constraint-based techniques allow efficient implementations of this analysis [14, 31]. We have developed a constraint-

[1]A flow-insensitive analysis ignores the flow of control between program points. A context-insensitive analysis does not distinguish between different invocations of a procedure.

based approach that extends the previous work with features necessary for points-to analysis for Java. We introduce *constraint annotations*, and show how to implement the analysis using annotated inclusion constraints of the form $L \subseteq_a R$, where a is a constraint annotation, and L and R are expressions representing points-to sets. The annotations play two roles in our analysis. *Method annotations* are used to model precisely and efficiently the semantics of virtual calls, by representing the relationships between a virtual call, its receiver objects, and its target methods. *Field annotations* allow separate tracking of the flow of values through the different fields of an object. By using techniques for efficient representation and resolution of systems of annotated inclusion constraints, we have been able to perform practical and precise points-to analysis for Java.

One disadvantage of Andersen's analysis is the implicit assumption that all code in the program is executable. Java programs contain large portions of unused library code; including such dead code can have negative effects on analysis cost and precision. In our analysis, we keep track of all methods potentially reachable from the entry points of the program, and we only analyze such reachable methods.

We have implemented our analysis and evaluated its performance on a large set of Java programs. On 16 out of the 23 data programs, analysis time is less than a minute. Even on large programs, the analysis runs in a few minutes and uses less than 180Mb of memory. Our results show that the analysis runs in practical time and space, which makes it a realistic candidate for a relatively precise general-purpose points-to analysis for Java.

We have evaluated the impact of the analysis on several of its possible client applications. Our results show very good analysis precision in determining which objects may be read or written by program statements; this *object read-write information* is a prerequisite for clients such as side-effect analysis, def-use analysis and dependence analysis. In addition, our measurements show significant improvement in the precision of the program call graph. Through profiling experiments, we observe that in many cases our analysis allows resolution of the majority of run-time virtual calls. Our experiments also show that the points-to solution can be used to detect a large number of objects that do not need synchronization or can be stack-allocated instead of heap-allocated.

Contributions. The contributions of our work are the following:

- We define a general-purpose points-to analysis for Java based on Andersen's points-to analysis for C. We show how to implement the analysis using a constraint-based approach that employs annotated inclusion constraints. The implementation models virtual calls and object fields precisely and efficiently, and only analyzes reachable methods.
- We evaluate the analysis on a large set of Java programs. Our results show that the analysis runs in practical time and space, and has significant impact on object read-write information, call graph construction, virtual call resolution, synchronization removal, and stack-based object allocation.

```
class Y {..}
class X {
  Y f;
  void set (Y r)
      { this.f = r; }
  static void main() {
   s1: X p = new X();
   s2: Y q = new Y();
      p.set(q);
  }
}
```

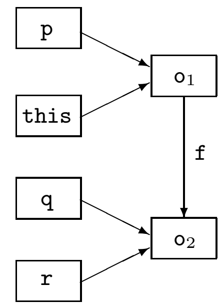


Figure 1: Sample program and its points-to graph.

Outline. The rest of the paper is organized as follows. Section 2 defines the semantics of our points-to analysis. Section 3 discusses the applications of points-to analysis for Java. Section 4 describes the general structure of our annotated inclusion constraints, and Section 5 contains the details of our constraint-based points-to analysis. The experimental results are presented in Section 6. Section 7 discusses related work, and Section 8 presents conclusions and future work.

2. SEMANTICS OF POINTS-TO ANALYSIS FOR JAVA

In this section we define the semantics of our points-to analysis for Java; Section 5 describes the implementation of the analysis with annotated inclusion constraints. The analysis is defined in terms of three sets. Set R contains all reference variables in the analyzed program (including static variables). Set O contains names for all objects created at object allocation sites; for each allocation site s_i, we use a separate object name $o_i \in O$. Set F contains all instance fields in program classes. Analysis semantics is expressed as manipulations of *points-to graphs* containing two kinds of edges. Edge $(r, o_i) \in R \times O$ shows that reference variable r points to object o_i. Edge $(\langle o_i, f\rangle, o_j) \in (O \times F) \times O$ shows that field f of object o_i points to object o_j. A sample program and its points-to graph are shown in Figure 1.

To simplify the presentation, we only discuss the kinds of statements listed below; our actual implementation (described in Section 5) handles the entire language.

- Direct assignment: `l = r`
- Instance field write: `l.f = r`
- Instance field read: `l = r.f`
- Object creation: `l = new C`
- Virtual invocation: $\texttt{l = r}_0\texttt{.m(r}_1\texttt{,...,r}_k\texttt{)}$

At a virtual call, name `m` uniquely identifies a method in the program. This method is the *compile-time* target of the call, and is determined based on the declared type of $\texttt{r}_0$ [18, Section 15.11.3]. At run time, the invoked method is determined by examining the class of the receiver object and all of its superclasses, and finding the first method that matches the signature and the return type of `m` [18, Section 15.11.4].

$$f(G, l = new\ C) = G \cup \{(l, o_i)\}$$

$$f(G, l = r) = G \cup \{(l, o_i) \mid o_i \in Pt(G, r)\}$$

$$f(G, l.f = r) = \\ G \cup \{(\langle o_i, f\rangle, o_j) \mid o_i \in Pt(G, l) \wedge o_j \in Pt(G, r)\}$$

$$f(G, l = r.f) = \\ G \cup \{(l, o_i) \mid o_j \in Pt(G, r) \wedge o_i \in Pt(G, \langle o_j, f\rangle)\}$$

$$f(G, l = r_0.m(r_1, \ldots, r_n)) = \\ G \cup \{resolve(G, m, o_i, r_1, \ldots, r_n, l) \mid o_i \in Pt(G, r_0)\}$$

$$resolve(G, m, o_i, r_1, \ldots, r_n, l) = \\ \texttt{let}\ m_j(p_0, p_1, \ldots, p_n, ret_j) = dispatch(o_i, m)\ \texttt{in} \\ \{(p_0, o_i)\} \cup f(G, p_1 = r_1) \cup \ldots \cup f(G, l = ret_j)$$

Figure 2: Points-to effects of program statements.

Analysis semantics is defined in terms of rules for adding new edges to points-to graphs. Each rule represents the semantics of a program statement. Figure 2 shows the rules as functions of the form $f : PtGraph \times Stmt \rightarrow PtGraph$. The points-to set (i.e., the set of all successors) of x in graph G is denoted by $Pt(G, x)$. The solution computed by the analysis is a points-to graph that is the closure of the empty graph under the edge-addition rules.

For most statements, the effects on the points-to graph are straightforward; for example, statement $l = r$ creates new points-to edges from l to all objects pointed to by r. For virtual call sites, resolution is performed for every receiver object pointed to by r_0. Function *dispatch* uses the class of the object and the compile-time target of the call to determine the actual method m_j invoked at run time. Variables $p_0, \ldots, p_n$ are the formal parameters of the method; variable p_0 corresponds to the implicit parameter `this`. Variable ret_j contains the return value of m_j.

3. APPLICATIONS OF POINTS-TO ANALYSIS FOR JAVA

Using points-to analysis in optimizing compilers and software engineering tools has several advantages. A single points-to analysis can enable a wide variety of client applications. The cost of the analysis can be amortized across many clients. Once implemented, the analysis can be reused by various clients at no additional development cost; such reusability is an important practical advantage. In this section we briefly discuss several specific applications of points-to analysis for Java. In our experiments, we have evaluated the impact of our analysis on some of these applications; the results from these experiments are presented in Section 6.

3.1 Object Read-Write Information

Points-to analysis can be used to determine which objects are read and/or written by every program statement. This information is a necessary prerequisite for a variety of other analyses. For example, for the purposes of side-effect analysis, points-to information can be used to answer questions like *"Can statement $p.f = x$ modify the f field of any object pointed to by q?"*. Points-to information is also needed to answer questions like *"Can statement $z = q.f$ read any memory locations that were written by statement $p.f = x$?"*, which are necessary for def-use analysis and dependence analysis.

Analyses that require read-write information are used in compilers to perform various optimizations such as code motion and partial redundancy elimination. In addition, such analyses play an important role in a variety of software engineering tools (e.g., in the context of program slicing or data-flow-based testing). Practical and precise points-to analysis is crucial for enabling the use of these analyses and optimizations.

3.2 Call Graph Construction and Virtual Call Resolution

The points-to solution can be used to determine the target methods of a virtual call by examining the classes of all possible receiver objects. The set of target methods is needed to construct the *call graph* for the analyzed program; this graph is a prerequisite for all interprocedural analyses and optimizations used in Java compilers and tools. If the call has only one target method, it can be *resolved* by replacing the virtual call with a direct call; this optimization eliminates the run-time overhead of virtual dispatch. In addition, virtual call resolution allows subsequent inlining of the target method, potentially enabling additional optimizations within the caller.

3.3 Synchronization Removal

Synchronization in Java allows safe access to shared objects in multi-threaded programs. Each object has an associated lock which is used to ensure mutual exclusion. Synchronization operations on locks can have considerable run-time overhead; this overhead occurs even in single-threaded programs, because the standard Java libraries are written in thread-safe manner.

Static analysis can be used to detect properties that allow the removal of unnecessary synchronization. For example, no synchronization is necessary for an object that cannot "escape" its creating thread and therefore cannot be accessed by any other thread (i.e., a *thread-local* object).[2] Some escape analyses [10, 7, 8, 36] have been used to identify thread-local objects and to remove the synchronization constructs associated with such objects.

Points-to analysis can be used as an alternative to escape analysis in detecting thread-local objects. Consider an object o_i and suppose that in the points-to graph computed by the analysis, o_i is not reachable from (i) static (i.e., global) reference variables, or (ii) objects of classes that implement interface `java.lang.Runnable`[3]. It can be proven that in this case o_i is not accessible outside the thread that created it. We can identify such thread-local objects by performing a reachability computation on the points-to graph; this approach is similar to the multithreaded object analysis proposed by Aldrich et al. [4].

[2]If synchronization operations are removed for objects receiving `wait`, `notify`, or `notifyAll` messages, the modified program may throw `IllegalMonitorStateException`. This problem can be avoided by maintaining the information needed by the notification methods without performing actual synchronization [28].

[3]The `run` methods of such objects are the starting points of new threads.

3.4 Stack Allocation

In some cases, an object can be allocated on a method's stack frame rather than on the heap. This transformation reduces garbage collection overhead and enables additional optimizations such as object reduction [17]. Similarly to synchronization removal, static analysis can be used to detect properties that allow stack-based allocation. For example, stack allocation is possible for an object that may never "escape" the lifetime of its creating method and therefore can only be accessed during that lifetime (i.e., a *method-local* object). Some escape analyses [10, 7, 36] can detect method-local objects; clearly, such objects can be allocated on the stack frames of their creating methods.

Points-to analysis can be used as an alternative to escape analysis in identifying method-local objects. Suppose that object o_i has been classified as thread-local according to the points-to solution (i.e., o_i is not reachable from static variables or from objects implementing `Runnable`). Also, suppose that in the computed points-to graph, o_i is not reachable from the formal parameters or the return variable of the method that created o_i. In this case, it can be proven that o_i is method-local; we can identify such method-local objects by traversing the points-to graph.

4. SYSTEMS OF ANNOTATED INCLUSION CONSTRAINTS

This section describes the general structure of the annotated inclusion constraints used in our points-to analysis for Java. The details about the specific kinds of constraints and annotations are discussed in Section 5.

Previous constraint-based implementations of Andersen's analysis for C [14, 31] employ non-annotated inclusion constraints. We have developed a constraint-based approach that extends this previous work by introducing constraint annotations. In our analysis, the annotations are used to model the flow of values between a virtual call site and the run-time target methods of the call. In addition, the annotations allow separate tracking of different object fields, which is not possible with the constraints from [14, 31].

4.1 Constraint Language

We consider annotated set-inclusion constraints of the form $L \subseteq_a R$, where a is chosen from a given set of annotations. We assume that one element of this set is designated as the empty annotation ϵ, and we use $L \subseteq R$ to denote constraints labeled with it. L and R are expressions representing sets, defined by the following grammar:

$$L, R \rightarrow v \mid c(v_1, \ldots, v_n) \mid proj(c, i, v) \mid 0 \mid 1$$

Here v and v_i are set variables, $c(\ldots)$ are constructed terms and $proj(\ldots)$ are projection terms. Each *constructed term* is built from an n-ary constructor c. A constructor is either *covariant* or *contravariant* in each of its arguments; the role of this variance in constraint resolution will be explained shortly. Constructed terms may appear on both sides of inclusion relations. 0 and 1 represent the empty set and the universal set; they are treated as nullary constructors. *Projections* of the form $proj(c, i, v)$ are terms used to select the i-th argument of any constructed term $c(.., v_i, ..)$, as described shortly. Projection terms may appear only on the right-hand side of an inclusion.

$$c(v_1, ..., v_n) \subseteq_a c(v_1', ..., v_n') \Rightarrow$$

$$\begin{cases} v_i \subseteq_a v_i' & \text{if } c \text{ is covariant in } i \text{ for } i = 1 \ldots n \\ v_i' \subseteq_a v_i & \text{if } c \text{ is contravariant in } i \text{ for } i = 1 \ldots n \end{cases}$$

$$c(v_1, ..., v_n) \subseteq_a proj(c, i, v) \Rightarrow$$

$$\begin{cases} v_i \subseteq_a v & \text{if } c \text{ is covariant in } i \\ v \subseteq_a v_i & \text{if } c \text{ is contravariant in } i \end{cases}$$

Figure 3: Resolution rules for non-atomic constraints.

4.2 Annotated Constraint Graphs

Systems of constraints from the above language can be represented as directed multi-graphs. Constraint $L \subseteq_a R$ is represented by an edge from the node for L to the node for R; the edge is labeled with the annotation a. There could be multiple edges between the same pair of nodes, each with a different annotation.

The nodes in the graph can be classified as variables, sources, and sinks. *Sources* are constructed terms that occur on the left-hand side of inclusions. *Sinks* are constructed terms or projections that occur on the right-hand side of inclusions. The graph only contains edges that represent *atomic constraints* of the following forms: $Source \subseteq_a Var$, $Var \subseteq_a Var$, or $Var \subseteq_a Sink$. If the constraint system contains a non-atomic constraint, the resolution rules from Figure 3 are used to generate new atomic constraints, as described in Section 4.3.

We use annotated constraint graphs based on the *inductive form* representation [3]. Inductive form is an efficient sparse representation that does not explicitly represent the transitive closure of the constraint graph. The graphs are represented with adjacency lists $pred(n)$ and $succ(n)$ stored at each node n. Edge (n_1, n_2, a), where a is an annotation, is represented either as a predecessor edge by having $\langle n_1, a\rangle \in pred(n_2)$, or as a successor edge by having $\langle n_2, a\rangle \in succ(n_1)$, but not both. $Source \subseteq_a Var$ is always a predecessor edge and $Var \subseteq_a Sink$ is always a successor edge. $Var \subseteq_a Var$ is either a predecessor or a successor edge, based on a fixed total order $\tau : Vars \rightarrow \mathcal{N}$. Edge (v_1, v_2, a) is a predecessor edge if and only if $\tau(v_1) < \tau(v_2)$. The order function is typically based on the order in which variables are created as part of building the constraint system [31].

4.3 Solving Systems of Annotated Constraints

Every system of annotated inclusion constraints can be represented by an annotated constraint graph in inductive form. The system is solved by computing the closure of the graph under the following transitive closure rule:

$$\left.\begin{array}{l} \langle L, a\rangle \in pred(v) \\ \langle R, b\rangle \in succ(v) \\ Match(a, b) \end{array}\right\} \Rightarrow L \subseteq_{a \circ b} R \qquad (\text{TRANS})$$

The closure rule can be applied locally, by examining $pred(v)$ and $succ(v)$. The new transitive constraint is created *only if* the annotations of the two existing constraints "match"—that is, only if $Match(a, b)$ holds, where $Match$ is a binary predicate on the set of annotations. Intuitively, the TRANS rule uses the annotations to filter out some flow of values in the constraint system. The $Match$ predicate is defined as follows:

$$Match(a, b) = \begin{cases} \text{true} & \text{if } a \text{ or } b \text{ is the empty annotation } \epsilon \\ \text{true} & \text{if } a = b \\ \text{false} & \text{otherwise} \end{cases}$$

The annotation of the new constraint is

$$a \circ b = \begin{cases} a & \text{if } b = \epsilon \\ b & \text{if } a = \epsilon \\ \epsilon & \text{otherwise} \end{cases}$$

Intuitively, an annotation is propagated until it is matched with another instance of itself, after which the two instances cancel out.

If the new constraint generated by the TRANS rule is atomic, a new edge is added to the graph. Otherwise, the resolution rules from Figure 3 are used to transform the constraint into several atomic constraints and their corresponding edges are added to the graph.

The closure of a constraint graph under the TRANS rule is the *solved inductive form* of the corresponding constraint system. The least solution of the system is not explicit in the solved inductive form [3], but is easy to compute by examining all predecessors of each variable. For constraint graphs without annotations, the least solution $LS(v)$ for a variable v is

$$LS(v) = \{c(\dots) \mid c(\dots) \in pred(v)\} \cup \bigcup_{u \in pred(v)} LS(u)$$

In this case, $LS(v)$ can be computed by transitive acyclic traversal of all predecessor edges [14]. For an annotated constraint graph, the traversal is done similarly, but the annotations are used as in rule TRANS:

$$\begin{aligned} LS(v) = \{\langle c(\dots), a\rangle \mid \langle c(\dots), a\rangle \in pred(v)\} \cup \\ \{\langle c(\dots), x \circ y\rangle \mid \langle u, x\rangle \in pred(v) \wedge \langle c(\dots), y\rangle \in LS(u) \\ \wedge\ Match(x, y)\} \end{aligned}$$

5. POINTS-TO ANALYSIS FOR JAVA USING ANNOTATED CONSTRAINTS

In this section we show how to implement the points-to analysis from Section 2 using annotated inclusion constraints. Recall that the analysis is defined in terms of the set R of all reference variables and the set O of names for all objects created at object allocation sites. Every element of $R \cup O$ is essentially an abstract memory location representing a set of run-time memory locations.

To implement our analysis with annotated inclusion constraints, we generalize an approach for modeling Andersen's analysis for C with non-annotated constraints [14, 31]. For each abstract location x, a set variable v_x represents the set of abstract locations pointed to by x. The representation of each location is through a ternary constructor *ref* which is used to build constructed terms of the form $ref(x, v_x, \overline{v_x})$. The last two arguments are the same variable, but with different variance—the overline notation is used to denote a contravariant argument. Intuitively, the second argument is

$$\langle l = new\ o_i\rangle \Rightarrow \{ref(o_i, v_{o_i}, \overline{v_{o_i}}) \subseteq v_l\}$$

$$\langle l = r\rangle \Rightarrow \{v_r \subseteq v_l\}$$

$$\langle l.f = r\rangle \Rightarrow \{v_l \subseteq proj(ref, 3, u),\ v_r \subseteq_f u\},\ u \text{ fresh}$$

$$\langle l = r.f\rangle \Rightarrow \{v_r \subseteq proj(ref, 2, u),\ u \subseteq_f v_l\},\ u \text{ fresh}$$

Figure 4: Constraints for assignment statements.

```
s1:  p = new X();
s2:  q = new Y();
     p.f = q;
     r = p.f;
```

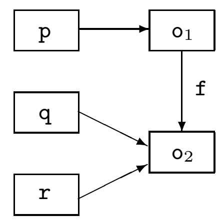


Figure 5: Accessing object fields.

used to read the values of locations pointed to by x, while the last argument is used to update the values of locations pointed to by x. Given a reference variable $r \in R$ and an object variable $o \in O$, constraint

$$ref(o, v_o, \overline{v_o}) \subseteq v_r$$

shows that r points to o.

We use *field annotations* to model the flow of values through fields of objects. Field annotations are unique identifiers for all instance fields defined in program classes. For any two object variables o_1 and o_2, constraint

$$ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq_f v_{o_1}$$

shows that field f in object o_1 points to object o_2.

5.1 Constraints for Assignment Statements

For every program statement, our analysis generates annotated inclusion constraints representing the semantics of the statement. Figure 4 shows the constraints generated for assignment statements. The first two generation rules are straightforward. The rule for $l.f = r$ uses the first constraint to access the points-to set of l, and the second constraint to update the values of field f in all objects pointed to by l. Similarly, the last rule uses two constraints to read the values of field f in all objects pointed to by r.

5.1.1 *Example*

Consider the statements in Figure 5 and their corresponding points-to graph. After processing the statements, our analysis creates the following constraints:

$$\begin{array}{ll} ref(o_1, v_{o_1}, \overline{v_{o_1}}) \subseteq v_p & ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_q \\ v_p \subseteq proj(ref, 3, u) & v_q \subseteq_f u \\ v_p \subseteq proj(ref, 2, w) & w \subseteq_f v_r \end{array}$$

where u and w are fresh variables. For the purpose of this example we assume that the variable order τ (defined in Section 4.2) is $\tau(v_p) < \tau(v_q) < \tau(v_r) < \tau(v_{o_1}) < \tau(v_{o_2}) <$

$\tau(u) < \tau(w)$. Consider the indirect write in `p.f = q`. Since we have

$$ref(o_1, v_{o_1}, \overline{v_{o_1}}) \subseteq v_p \subseteq proj(ref, 3, u)$$

we can use the TRANS rule and the resolution rules from Figure 3 to generate a new constraint $u \subseteq v_{o_1}$. Thus,

$$v_q \subseteq_f u \subseteq v_{o_1}$$

and using rule TRANS we generate $v_q \subseteq_f v_{o_1}$. Intuitively, this new constraint shows that some of the values of field f in object o_1 come from variable q. Now we have

$$ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_q \subseteq_f v_{o_1}$$

Since both constraint edges are predecessor edges, we cannot apply rule TRANS. Still, in the least solution of the constraint system (as defined in Section 4.3), we have the constraint $ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq_f v_{o_1}$, which shows that field f of o_1 points to o_2.

To model indirect reads, we use the second argument of the *ref* constructor. For example, for the constraints above we have

$$ref(o_1, v_{o_1}, \overline{v_{o_1}}) \subseteq v_p \subseteq proj(ref, 2, w)$$

and therefore $v_{o_1} \subseteq w \subseteq_f v_r$, which through TRANS generates $v_{o_1} \subseteq_f v_r$. This new constraint shows that the value of r comes from field f of object o_1. Now we have

$$v_q \subseteq_f v_{o_1} \subseteq_f v_r$$

Since the annotations of the two constraints match—that is, they represent accesses to the same field—we generate $v_q \subseteq v_r$ to represent the flow of values from q to r. Thus, in the least solution of the system we have

$$ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_r$$

which shows that reference variable r points to o_2. This example illustrates how field annotations allow us to model the flow of values through object fields.

5.2 Handling of Virtual Calls

For every virtual call in the program, our analysis generates a constraint according to the following rule:

$$\langle l = r_0.m(r_1, \ldots, r_k)\rangle \Rightarrow$$
$$\{v_{r_0} \subseteq_m lam(\overline{0}, \overline{v_{r_1}}, \ldots, \overline{v_{r_k}}, v_l)\}$$

The rule is based on a *lam* (lambda) constructor. The constructor is used to build a term that encapsulates the actual arguments and the left-hand-side variable of the call. The annotation on the constraint is a unique identifier of the *compile-time* target method of the call. This annotation is used during the analysis to find all appropriate *run-time* target methods.

To model the semantics of virtual calls as defined in Section 2, we separately perform virtual dispatch for every receiver object pointed to by r_0. In order to do this efficiently, we use a precomputed *lookup table*. For a given receiver object at a virtual call site, the lookup table is used to determine the corresponding run-time target method, based on the class of the receiver object.[4] Such a table is straightforward to precompute by analyzing the class hierarchy; the table is essentially a representation of the *dispatch* function from Section 2.

Given the class of the receiver object and the unique identifier for the compile-time target of the virtual call, the lookup table returns a lambda term of the form

$$lam(\overline{v_{p_0}}, \overline{v_{p_1}}, \ldots, \overline{v_{p_k}}, v_{ret})$$

Here p_i are the formal parameters of the run-time target method; p_0 corresponds to the implicit parameter `this`. We assume that each method has a unique variable *ret* that is assigned the value returned by the method (this can be achieved by inserting auxiliary assignments in the program representation). At the beginning of the analysis, lambda terms of the above form are created for all non-abstract methods in the program and are stored in the lookup table.

To model the effects of virtual calls, we define an additional closure rule VIRTUAL. This rule encodes the semantics of virtual calls described in Section 2 and is used together with the TRANS rule to obtain the solved form of the constraint system. VIRTUAL is applied whenever we have two constraints of the form

$$ref(o, v_o, \overline{v_o}) \subseteq v \qquad v \subseteq_m lam(\overline{0}, \overline{v_{r_1}}, \ldots, \overline{v_{r_k}}, v_l)$$

As described in Section 4.2, the edge from the *ref* term is a predecessor edge, and the edge to the *lam* term is a successor edge. Thus, the VIRTUAL closure rule can be applied locally, by examining sets $pred(v)$ and $succ(v)$. Whenever two such constraints are detected, the lookup table is used to find the lambda term for the run-time method corresponding to object o and compile-time target method m. The result of applying VIRTUAL are two new constraints:

$$\begin{aligned} ref(o, v_o, \overline{v_o}) &\subseteq v_{p_0} \\ lam(\overline{v_{p_0}}, \overline{v_{p_1}}, \ldots, \overline{v_{p_k}}, v_{ret}) &\subseteq lam(\overline{0}, \overline{v_{r_1}}, \ldots, \overline{v_{r_k}}, v_l) \end{aligned}$$

The first constraint creates the association between parameter `this` of the invoked method and the receiver object. The second constraint immediately resolves to $v_{r_i} \subseteq v_{p_i}$ (for $i \geq 1$) and $v_{ret} \subseteq v_l$, plus the trivial constraint $0 \subseteq v_{p_0}$. These new atomic constraints model the flow of values from actuals to formals, as well as the flow of return values to the left-hand side variable l used at the call site.

5.2.1 *Example*

Consider the set of statements in Figure 6. For the purpose of this example, assume that $\tau(v_a) < \tau(v_b) < \tau(v_c)$. Since the declared type of `b` is `B`, at call site c_1 the compile-time target method is `B.n`; thus, we have

$$v_b \subseteq_{B.n} lam(\overline{0}, v_x)$$

When rule VIRTUAL is applied as shown in (1), the lookup for receiver object o_2 and compile-time target `B.n` produces run-time target `B.n`. The resolution with the *lam* term for `B.n` creates the two new constraints shown in (1).

The declared type of `c` is `A`, and for call site c_2 we have $v_c \subseteq_{A.n} lam(\overline{0}, v_y)$. Thus,

$$ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_b \qquad v_b \subseteq_{A.n} lam(\overline{0}, v_y)$$

where the second constraint is obtained through the TRANS

[4] Every object is tagged with its class; this tag is used when performing lookups.

```
class A { X n() { ...  return rA; } }
class B extends A
        { X n() { ...  return rB; } }
s1:  A a = new A();
s2:  B b = new B();
     A c = b;
c1:  X x = b.n();
c2:  X y = c.n();
     if (...)  a = b;
c3:  X z = a.n();
```

(1) $ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_b \subseteq_{B.n} lam(\overline{0}, v_x) \Rightarrow$

$\{ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_{B.n.this},\ v_{rB} \subseteq v_x\}$

(2) $ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_b \subseteq_{A.n} lam(\overline{0}, v_y) \Rightarrow$

$\{ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_{B.n.this},\ v_{rB} \subseteq v_y\}$

(3) $ref(o_1, v_{o_1}, \overline{v_{o_1}}) \subseteq v_a \subseteq_{A.n} lam(\overline{0}, v_z) \Rightarrow$

$\{ref(o_1, v_{o_1}, \overline{v_{o_1}}) \subseteq v_{A.n.this},\ v_{rA} \subseteq v_z\}$

(4) $ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_a \subseteq_{A.n} lam(\overline{0}, v_z) \Rightarrow$

$\{ref(o_2, v_{o_2}, \overline{v_{o_2}}) \subseteq v_{B.n.this},\ v_{rB} \subseteq v_z\}$

Figure 6: Example of virtual call resolution.

rule.[5] When applying rule VIRTUAL as shown in (2), the lookup for receiver object o_2 and compile-time target `A.n` leads to run-time target `B.n`. The two new constraints which result from the resolution are shown in (2). For call site c_3, the receiver object can be either o_1 or o_2. As shown in (3) and (4), separate lookup and resolution is performed for each receiver.

5.3 Correctness

For every program statement, our analysis generates constraints representing the semantics of the statement. This initial constraint system is solved by closing the corresponding constraint graph under closure rules TRANS and VIRTUAL. Let A^* be the solved inductive form of the constraint system. Recall that the least solution of the system is not explicit in A^* and can be obtained through additional traversal of predecessor edges, as described in Section 4.3.

Let G^* be the points-to graph computed by the algorithm in Section 2. Consider a reference variable r and an object variable o such that $(r, o) \in G^*$. It can be proven that the least solution constructed from A^* contains the constraint $ref(o, v_o, \overline{v_o}) \subseteq v_r$. Similarly, consider two object variables o_i and o_j such that $(\langle o_i, f\rangle, o_j) \in G^*$; it can be proven that the least solution contains $ref(o_j, v_{o_j}, \overline{v_{o_j}}) \subseteq_f v_{o_i}$.

The proof of the these claims depends on the following restriction on the variable order τ: all variables v_r, where $r \in R$, should have lower order than the rest of the constraint variables. We enforce this restriction as part of building the constraint system. Given this restriction, it can be proven that the least solution of the constraint system represents all points-to pairs from G^* [26].

[5]Note that if $\tau(v_c) < \tau(v_b)$, instead of propagating the *lam* term to v_b we would propagate the *ref* term to v_c.

5.4 Cycle Elimination and Projection Merging

Cycle elimination [14] and projection merging [31] are two techniques that can be used to reduce the cost of Andersen's analysis for C. We have adapted these techniques to allow us to reduce the cost of our points-to analysis for Java.

The idea behind cycle elimination is to detect a set of variables that form a cycle in the constraint graph:

$$v_1 \subseteq v_2 \subseteq \ldots \subseteq v_k \subseteq v_1$$

Clearly, all such variables have equal solutions and can be replaced with a single variable. Whenever a cycle is detected during the resolution process, one variable from the cycle is chosen as a witness variable, and the rest of the variables are redirected to the witness. This transformation has no effect on the computed solution, but can significantly reduce the cost of the analysis.

Cycle detection is performed every time a new edge is added between two variables v_i and v_j. The detection algorithm essentially performs depth-first traversal of the constraint graph and tries to determine whether v_i is reachable from v_j. Cycle detection is partial and does not detect all cycles. Nevertheless, for Andersen's analysis for C this technique has significant impact on the running time of the analysis [14].

Cycle elimination cannot be used directly for the annotated constraint systems presented in this paper. If we performed the standard cycle detection, we would discover cycles in which some edges have field annotations; however, the variables in such cycles do not have the same solution, and cannot be replaced by a single witness variable. To guarantee the correctness of our analysis for Java, we use a restricted form of cycle elimination. The cycle detection algorithm is invoked only when a new edge is added between two reference variables—that is, when the new edge is (v_{r_i}, v_{r_j}), where $r_i, r_j \in R$. It can be proven that in this case, the detected cycle contains only reference variables, and all edges in the cycle have empty annotations. This guarantees that all variables on the cycle have identical points-to sets, and therefore replacing the cycle with a single variable preserves the points-to solution.

Projection merging is a technique for reducing redundant edge additions in constraint systems [31]. It combines multiple projection constraints for the same variable into a single projection constraint. For example, constraints $v \subseteq proj(c, i, u_1)$ and $v \subseteq proj(c, i, u_2)$ are replaced by

$$v \subseteq proj(c, i, w) \qquad w \subseteq u_1 \qquad w \subseteq u_2$$

where w is a special projection variable. For points-to analysis for C, constraints of the form $w \subseteq u_i$ are represented only as successor edges; this restriction guarantees a bound on the number of projection variables w. The analysis ensures the restriction by assigning to w a high index in the variable order τ [31]. In this case, projection merging is beneficial because it is coupled with cycle elimination.

In our annotated constraint systems, projection merging does not interact with cycle elimination. In our case, the high indices from [31] (which necessitate the interaction) are not required. The high indices become unnecessary because the bound on the number of special projection variables is ensured by the variable ordering restriction from Section 5.3. Thus, the special projection variables can be treated similarly to the rest of the variables in the constraint system. This *decoupled* form of projection merging has significant impact on the running time of the analysis for Java.

5.5 Tracking Reachable Methods

Andersen's analysis implicitly assumes that all code in the program is executable. Since Java programs heavily use libraries that contain many unused methods, we have augmented our analysis to keep track of reachable methods, in order to avoid analyzing dead code. Thus, we take into account the effects of statements in a method body only if the method has been shown to be reachable from one of the entry methods of the program. The set of entry methods contains (i) the `main` method of the starting class, (ii) the methods invoked at JVM startup (e.g., `initializeSystemClass`), and (iii) the class initialization methods `<clinit>` containing the initializers for static fields [22, Section 3.9].

During the analysis, we maintain a list of reachable methods; whenever a method becomes reachable, all statements in its body are processed and the appropriate constraints are introduced in the constraint system. Any call to a constructor also generates a corresponding call to the appropriate `finalize` method. For multi-threaded programs, a call to `Thread.start` is treated as a call to the corresponding `run` method.

5.6 Analysis Implementation

We use the Soot framework (`www.sable.mcgill.ca`), version 1.0.0, to process Java bytecode and to build a typed intermediate representation [35]. The constraint-based analysis uses BANE (Berkeley ANalysis Engine) [2]. BANE is a toolkit for constructing constraint-based program analyses. The public distribution of BANE (`bane.cs.berkeley.edu`) contains a constraint-solving engine for non-annotated constraints that employs inductive form, cycle elimination, and projection merging. We modified the constraint engine to represent and solve systems of annotated constraints. The analysis works on top of the constraint engine, by processing newly discovered reachable methods and generating the appropriate constraints. The points-to effects of JVM startup code and native methods (for JDK 1.1.8) are encoded in stubs included in the analysis input. Dynamic class loading (e.g., through `Class.forName`) and reflection (e.g., calls to `Class.newInstance`) are resolved manually; similar approaches are typical for static whole-program compilers and tools [20, 16, 33, 34].

6. EMPIRICAL RESULTS

All experiments were performed on a 360MHz Sun Ultra-60 machine with 512Mb physical memory. The reported times are the median values out of three runs. We used 23 publicly available data programs, ranging in size from 56Kb to about 1Mb of bytecode. We used programs from the SPEC JVM98 suite, other benchmarks used in previous work on analysis for Java, as well as programs from an Internet archive (`www.jars.com`) of popular publicly available Java applications.

Table 1 shows some characteristics of the data programs.

Program	User Class	Size (Kb)	Whole-program Class	Method	Stmt
proxy	18	56.6	565	3283	58837
compress	22	76.7	568	3316	60010
db	14	70.7	565	3339	60747
jb-6.1	21	55.6	574	3393	60898
echo	17	66.7	577	3544	62646
raytrace	35	115.9	582	3451	62755
mtrt	35	115.9	582	3451	62760
jtar-1.21	64	185.2	618	3583	65112
jlex-1.2.5	25	95.1	578	3381	65437
javacup-0.10	33	127.3	581	3564	66463
rabbit-2	52	157.4	615	3770	68277
jack	67	191.5	613	3573	69249
jflex-1.2.2	54	198.2	608	3692	71198
jess	160	454.2	715	3973	71207
mpegaudio	62	176.8	608	3531	71712
jjtree-1.0	72	272.0	620	4078	79587
sablecc-2.9	312	532.4	864	5151	82418
javac	182	614.7	730	4470	82947
creature	65	259.7	626	3881	83454
mindterm1.1.5	120	461.1	686	4420	90451
soot-1.beta.4	677	1070.4	1214	5669	92521
muffin-0.9.2	245	655.2	824	5253	94030
javacc-1.0	63	502.6	615	4198	102986

Table 1: Characteristics of the data programs. First two columns show the number and bytecode size of user classes. Last three columns include library classes.

The first two columns show the number of user (i.e., non-library) classes and their bytecode size. The next three columns show the size of the program, including library classes, after using class hierarchy analysis (CHA) [11] to filter out irrelevant classes and methods.[6] The number of methods is essentially the number of nodes in the call graph computed by CHA. The last column shows the number of statements in Soot's intermediate representation.

6.1 Analysis Cost

Our first set of experiments measured the cost of the analysis, as shown in Table 2. The first two columns show the running time of the analysis and the amount of memory used. For 16 out of the 23 programs, the analysis runs in less than a minute. For all programs, the running time is less than six minutes and the memory usage is less than 180Mb. These results show that our analysis is practical in terms of running time and memory usage, as evidenced on a large set of Java programs. This practicality means that the analysis can be used as a relatively precise general-purpose points-to analysis for advanced static compilers and software engineering tools for Java.

Analysis cost can be reduced further if the library code is analyzed in advance. This would allow certain partial analysis information about the Java libraries to be computed once and subsequently used multiple times for different client programs. We intend to investigate this approach in our future work.

We also investigated a version of our analysis in which no

[6]CHA is an inexpensive analysis that determines the possible targets of a virtual call by examining the class hierarchy of the program.

Program	Time (sec)	Memory (Mb)	Time-nf (sec)	Memory-nf (Mb)
proxy	6.5	38.9	29.8	45.1
compress	22.2	45.3	68.4	77.2
db	23.2	46.8	63.2	80.7
jb	13.3	40.7	28.1	45.6
echo	41.5	61.8	231.6	184.3
raytrace	26.1	51.2	99.6	100.4
mtrt	26.1	49.0	89.7	100.2
jtar	44.7	58.8	125.8	113.3
jlex	17.7	45.7	137.7	83.6
javacup	32.0	54.1	80.5	83.8
rabbit	27.9	53.3	74.3	82.7
jack	48.7	63.1	5871.4	134.8
jflex	56.1	73.1	208.6	191.8
jess	41.8	64.9	202.7	207.7
mpegaudio	28.1	51.8	227.1	226.3
jjtree	24.6	52.5	57.4	76.0
sablecc	287.2	151.9	815.4	280.9
javac	350.0	151.5	396.8	334.1
creature	176.4	101.0	821.4	319.7
mindterm	94.6	95.4	407.7	341.0
soot	239.5	176.4	401.8	308.5
muffin	243.7	163.8	–	–
javacc	190.5	125.5	167.8	167.7

Table 2: Running time and memory usage of the analysis (with and without field annotations).

Program	1	2	3	4–5	6–9	≥10
proxy	53%	18%	8%	7%	8%	6%
compress	57%	12%	12%	6%	5%	8%
db	56%	12%	14%	7%	6%	5%
jb	59%	20%	7%	4%	5%	5%
echo	54%	15%	9%	8%	4%	10%
raytrace	51%	13%	12%	6%	10%	8%
mtrt	51%	13%	12%	6%	10%	8%
jtart	47%	13%	8%	10%	9%	13%
jlex	89%	4%	2%	2%	2%	1%
javacup	68%	11%	5%	3%	7%	6%
rabbit	48%	24%	11%	5%	6%	6%
jack	58%	10%	8%	5%	4%	15%
jflex	61%	12%	10%	4%	2%	11%
jess	48%	14%	14%	13%	5%	6%
mpegaudio	56%	15%	14%	6%	4%	5%
jjtree	54%	20%	8%	3%	10%	5%
sablecc	69%	13%	3%	3%	2%	10%
javac	49%	13%	9%	9%	4%	16%
creature	36%	50%	1%	1%	2%	10%
mindterm	69%	6%	8%	7%	2%	8%
soot	69%	12%	2%	10%	3%	4%
muffin	60%	16%	5%	5%	4%	10%
javacc	73%	9%	4%	3%	7%	4%

Table 3: Number of accessed objects for indirect access expressions. Each column shows the percentage of indirect accesses with a given number of objects.

field annotations are used, and therefore individual object fields are not distinguished. The last two columns in Table 2 show the cost of this *no-fields* version. The running time is between 88% and 12056% (average 892%, median 384%) of the running time of the original analysis; the memory usage is between 112% and 437% (215% on average). Typically, the *no-fields* version is significantly more expensive; for one of the larger programs, it even ran out of memory. These results show the importance of distinguishing object fields: the improved precision produces smaller points-to sets, which in turn reduces analysis cost. By using field annotations, we have been able to distinguish object fields in a simple and efficient manner.

6.2 Object Read-Write Information

We performed measurements to estimate the potential impact of our analysis on clients of object read-write information (e.g., side-effect analysis and def-use analysis). In particular, we considered all expressions of the form $p.f$ occuring in statements in reachable methods. For each such *indirect access expression*, the points-to set of p contains all objects that may be read or written by the corresponding statement. More precise points-to analyses produce smaller numbers of accessed objects; this improves the precision and reduces the cost of the clients of the read-write information. Thus, to estimate the potential impact of our analysis, we measured the number of accessed objects for each indirect access expression; similar metrics have been traditionally used for points-to analysis for C.

Table 3 shows the distribution of the number of accessed objects; each column corresponds to a specific range of numbers. For example, the first column corresponds to expressions that may only access a single object, while the last column corresponds to expressions that may access 10 or more objects. Each column shows what percentage of all indirect access expressions corresponds to the particular range of numbers of accessed objects.

The measurements in Table 3 indicate that our analysis produces precise read-write information. Typically, more than half of the indirect accesses are resolved to a single object (which is the lower bound for this metric), and on average 81% of the accesses are resolved to at most three objects. These results show that the analysis is a promising candidate for producing useful read-write information for clients such as (i) aggressive optimizing compilers, in which optimizations require precise read-write information, and (ii) software engineering tools, in which analysis precision is important for reducing the human effort spent on program understanding, restructuring, and testing.

6.3 Call Graph Construction and Virtual Call Resolution

To measure the precision with respect to call graph construction and virtual call resolution, we compared our points-to analysis with Rapid Type Analysis (RTA) [6]. RTA is an inexpensive and widely used analysis for call graph construction. It performs a reachability computation on the call graph generated by CHA; by keeping track of the classes that have been instantiated, RTA computes a more precise call graph than CHA.

Both our analysis and RTA improve the call graph computed by CHA by identifying sets of methods reachable from the entry points of the program; this reachability computation reduces the number of nodes in the call graph. For brevity, we summarize this reduction without explicitly showing the number of nodes for each program. The average reduction in the number of nodes is 54% for our analysis and 47% for RTA. On average, the call graph computed by

Program	(a) Removed Targets		(b) Resolved Call Sites	
	Points-to	RTA	Points-to	RTA
proxy	8.7	5.8	50%	9%
compress	6.4	2.4	58%	18%
db	6.0	2.2	61%	18%
jb	8.5	5.8	56%	13%
echo	3.4	1.4	45%	19%
raytrace	6.0	2.3	58%	21%
mtrt	6.0	2.3	58%	21%
jtar	5.3	3.2	39%	17%
jlex	9.3	6.2	63%	12%
javacup	6.3	3.9	63%	14%
rabbit	8.6	4.1	54%	14%
jack	3.0	0.9	83%	12%
jflex	6.5	3.5	45%	12%
jess	5.7	2.0	57%	15%
mpegaudio	7.1	2.2	53%	17%
jjtree	8.4	5.9	54%	14%
sablecc	7.1	1.5	32%	7%
javac	2.8	1.1	32%	15%
creature	3.8	2.3	50%	25%
mindterm	2.5	1.3	43%	24%
soot	4.5	1.0	41%	1%
muffin	5.7	2.0	48%	17%
javacc	5.0	2.8	79%	10%
Average	5.9	2.9	53%	15%

Table 4: Improvements for CHA-unresolved virtual call sites. (a) Average reduction in the number of target methods per call site. (b) Percentage of uniquely resolved call sites.

our analysis has 14% less nodes than the call graph computed by RTA. This reduction allows subsequent analyses and optimizations to safely ignore portions of the program.

To determine the improvement for call graph edges, we considered call sites that could not be resolved to a single target method by CHA. Let V be the set of all CHA-unresolved call sites that occur in methods identified by our analysis as reachable. For our data programs, the size of V is between 7% and 44% (22% on average) of all virtual call sites in reachable methods. For each site from V, we computed the difference between the number of target methods according to CHA and the number of target methods according to RTA and our analysis. The average differences are shown in the first section of Table 4. On average, our analysis removes more than twice as many targets as RTA; this improved precision is beneficial for reducing the cost and improving the precision of subsequent interprocedural analyses.

The second section of Table 4 shows the percentage of call sites from V that were resolved to a single target method. Our points-to analysis performs significantly better than RTA—on average, 53% versus 15% of the virtual call sites are resolved. The increased precision allows better removal of run-time virtual dispatch and additional method inlining.

We performed additional experiments to estimate the potential performance impact of analysis precision on virtual call resolution. These experiments used a subset of our data programs for which we had representative input data. For each program, we instrumented the user classes (i.e., non-library classes) and measured the number of times each call site was executed during a profile run of the program. Column (a) in Table 5 shows the total number of invocations of CHA-unresolved call sites. This number is an indicator of the run-time overhead of virtual dispatch, as well as the missed opportunities for performance improvement through inlining. We also measured what percentage of this total number was contributed by call sites that are uniquely resolved by RTA and by our analysis. These percentages are shown in column (b) in Table 5; higher percentages indicate higher potential for performance improvement.

Program	(a) Num Calls	(b) Resolved		(c) Run-time Monomorphic
		RTA	Points-to	
compress	0	—	—	—
db	10550782	0%	100%	100%
mtrt	2833913	0%	0%	0%
jlex	1336	11.0%	99.9%	100%
jack	2663305	5.9%	98.6%	98.6%
jess	1511933	0.03%	0.3%	50.2%
mpegaudio	6528736	0%	0%	0.5%
sablecc	1005390	0.1%	36.5%	46.7%
javac	20198864	0%	19.2%	71.0%
javacc	44322	0.02%	85.8%	85.9%

Table 5: Execution counts for virtual call sites. (a) Total count for CHA-unresolved call sites. (b) Percentage due to resolved sites. (c) Percentage due to sites with a single run-time target.

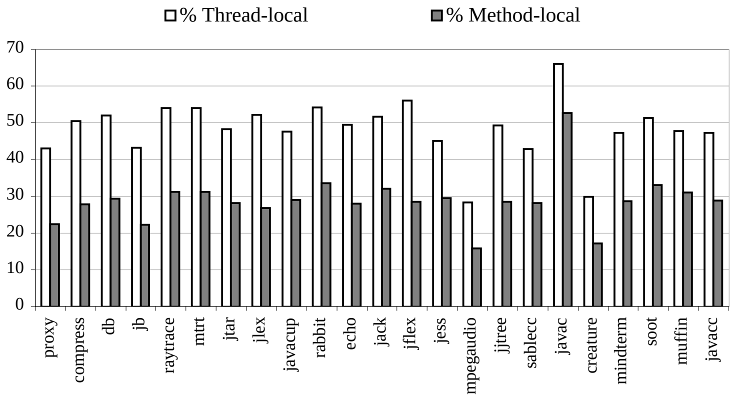
The results from this profiling experiment indicate that RTA has little potential for improving the run-time performance over CHA. Our analysis shows significantly higher potential, and for several programs it allows the resolution of the majority of run-time virtual calls. In addition, we used the profile to determine what CHA-unresolved call sites had only one run-time target. Column (c) shows the contribution of such sites to the total count from column (a). This number is an upper bound on the number of invocations that could be resolved by a static analysis. By comparing columns (b) and (c), it is clear that in many cases our analysis achieves performance close to the best possible performance.

6.4 Synchronization Removal and Stack Allocation

Points-to analysis has a wide variety of client applications, including optimizations for removal of unnecessary synchronization and for stack-based object allocation. To investigate the impact of our analysis on synchronization removal and stack allocation, we identified all object allocation sites that correspond to thread-local and method-local objects, as described in Section 3. Figure 7(a) shows what percentage of all allocation sites in reachable methods were identified as thread-local or method-local.

Across all programs, the analysis detects a significant number of allocation sites for thread-local objects—on average, about 48% of all allocation sites. These results indicate that the points-to information can be useful in detecting and eliminating unnecessary synchronization in Java programs. The analysis also discovers a significant number of sites for method-local objects—on average, about 29% of all sites. These results suggest that there are many opportunities for stack-based object allocation that can be detected with our analysis.

(a) Object allocation sites



(b) Run-time objects

Program	Objects	Thread-local	Method-local
compress	456	99.3%	39.0%
db	154325	0.03%	0.01%
mtrt	6457298	99.9%	85.0%
jlex	7350	50.9%	31.6%
jack	1340919	86.7%	77.0%
jess	7902221	17.9%	17.9%
mpegaudio	2025	12.4%	12.4%
sablecc	420494	24.9%	13.7%
javac	3738777	27.6%	21.2%
javacc	43265	65.7%	45.8%

Figure 7: (a) Thread-local and method-local allocation sites. (b) Number of objects created at run time: total number, percentage of thread-local objects, and percentage of method-local objects.

As with virtual call resolution, we performed additional profiling experiments to obtain better estimates of the potential impact on run-time performance. Using the same set of programs with instrumented user classes and the same data input sets, we measured the number of run-time objects created at each object allocation site. The total number of created objects is shown in the first column of Figure 7(b); the other two columns show what percentage of these objects were identified by our analysis as thread-local or method-local.

The results from this experiment indicate that our analysis has good potential for improving the run-time performance through synchronization removal and stack-based object allocation. The results in Figure 7(b) are similar to the results obtained through more expensive flow- and context-sensitive escape analyses [10, 36]. Even though direct comparison with this previous work is not possible (due to differences in the infrastructure and the data programs), the results suggest that our analysis may be a viable alternative to these more expensive analyses.

7. RELATED WORK

Points-to analysis for object references in Java is clearly related to pointer analysis for imperative languages such as C. There are various pointer analyses for C with different tradeoffs between cost and precision. The closest related work from this category are the constraint-based implementations of Andersen's analysis from [14, 31], in which non-annotated constraints are used together with inductive form, cycle elimination, and projection merging. We extend this work by introducing constraint annotations and by changing the constraint representation and the resolution procedure to allow points-to analysis for Java. Field annotations are used to track object fields separately; this is not possible with the constraints from [14, 31]. Method annotations allow us to model the semantics of virtual calls. In addition, we avoid analyzing dead library code by including a reachability computation in the analysis.

Constraint indices and constraint polarities [15] have been used to introduce context-sensitivity in unification-based flow analysis. This work has similar flavor to our use of annotations for tracking flow of values through object fields. Conceptually, in both cases the goal is to restrict the flow of values in constraint systems—either for unification-based constraints in [15], or for inclusion constraints in our case.

Recent work [30], which postdates our initial report [27], describes a points-to analysis for Java based on Andersen's analysis for C. Analysis cost is higher than ours, which is most likely due to the different kind of constraints employed by this approach. Another recent points-to analysis for Java based on Andersen's analysis is presented in [21], together with several analysis variations. Direct comparison with this work is not possible because it handles the library code in a different manner; based on the size of the analyzed code, our analysis appears to be faster. Another example of points-to analysis for object-oriented languages is due to Chatterjee et al. [9]. This flow- and context-sensitive analysis is more precise and more expensive than ours. Points-to analyses with different degrees of precision have been proposed in the context of a framework for call graph construction in object-oriented languages [19]. The closest to our work is the 1-1-CFA algorithm, which incorporates a flow- and context- sensitive points-to analysis. The scalability of the analyses from [19, 9] remains unclear; our approach may be a practical alternative to these more expensive analyses. Other related analyses, based on unification techniques, are a context-sensitive alias analysis for synchronization removal due to Ruf [28], and points-to analyses for Java [25, 30, 21] derived from Steensgaard's points-to analysis for C [29].

Class analysis for object-oriented languages computes a set of classes for each program variable; this set approximates the classes of all run-time values for this variable. The traditional client applications of class analysis are call graph construction and virtual call resolution. DeFouw et al. [12] present a family of practical interprocedural class analyses, ranging from linear to cubic complexity; the closest to our analysis are the classic 0-CFA and the linear-edge 0-CFA algorithms. Other work in this area considers more expensive analyses with some degree of context- or flow-sensitivity [23, 1, 24, 13, 19], as well as less precise but inexpensive analyses such as RTA [6, 34, 32]. Comparison with the results from this work is difficult due to differences in language, infrastructure, and analysis parameters (e.g., handling of libraries, dynamic class loading, etc.).

There is a large body of work on synchronization removal and stack-based object allocation [4, 10, 7, 8, 36, 17, 28]. Gay and Steensgaard [17] present a unification-based analysis for stack allocation. Ruf [28] describes an unification-based algorithm for synchronization removal. Aldrich et al. [4] propose several approaches for synchronization removal. Our approach for identifying thread-local objects is similar to the multithreaded object analysis from [4] (which is based on 1-1-CFA). Previous work on escape analysis for Java [10, 7, 8, 36] also investigates synchronization removal and stack allocation; the scalability of these approaches remains unclear. In contrast to these specialized analyses, we propose a points-to analysis that can also be used for a variety of other client applications.

8. CONCLUSIONS AND FUTURE WORK

Designing precise and practical points-to analyses is important for enabling a wide variety of popular analyses and optimizations. We define a points-to analysis for Java based on Andersen's points-to analysis for C. We implement the analysis by using a constraint-based approach which employs constraint annotations. Method annotations are used to model precisely and efficiently the semantics of virtual calls. Field annotations allow us to distinguish between different fields of an object. On a large set of Java programs, our experiments show that the cost of the analysis is practical. We also show that the points-to solution has significant impact on object read-write information, call graph construction, virtual call resolution, synchronization removal, and stack-based object allocation. Our results demonstrate that the analysis is a realistic candidate for a relatively precise, practical, general-purpose points-to analysis for advanced optimizing compilers and software engineering tools for Java.

One direction of future work is to investigate techniques for further reduction of analysis cost. For example, the cost can be reduced if the library code is analyzed in advance. This would allow partial analysis information about the Java libraries to be computed once and subsequently used for different client programs.

Another direction of future work is to investigate the impact of the analysis solution on traditional client analyses such as def-use analysis, side-effect analysis, and dependence analysis, which in turn are necessary for various optimizations. Such analyses and optimizations have been extensively investigated in other languages, and will play an increasingly important role in aggressive optimizing compilers for Java.

Finally, it would be interesting to investigate applications of points-to analysis in the context of software engineering tools for program checking, understanding, maintenance, and testing. The functionality provided by such tools will be necessary for the development of production-strength Java software systems.

9. ACKNOWLEDGMENTS

We would like to thank Matthew Arnold, Michael Hind, and the anonymous reviewers for their suggestions for improving the content and the presentation of this work. This research was supported by NSF grant CCR-9900988.

10. REFERENCES

[1] O. Agesen. Constraint-based type inference and parametric polymorphism. In *Static Analysis Symposium*, LNCS 864, pages 78–100, 1994.

[2] A. Aiken, M. Fähndrich, J. Foster, and Z. Su. A toolkit for constructing type- and constraint-based program analyses. In *International Workshop on Types in Compilation*, 1998.

[3] A. Aiken and E. Wimmers. Type inclusion constraints and type inference. In *Conference on Functional Programming Languages and Computer Architecture*, pages 31–41, June 1993.

[4] J. Aldrich, C. Chambers, E. G. Sirer, and S. Eggers. Static analyses for eliminating unnecessary synchronization from Java programs. In *Static Analysis Symposium*, LNCS 1694, pages 19–38, 1999.

[5] L. Andersen. *Program Analysis and Specialization for the C Programming Language*. PhD thesis, DIKU, University of Copenhagen, 1994.

[6] D. Bacon and P. Sweeney. Fast static analysis of C++ virtual function calls. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 324–341, 1996.

[7] B. Blanchet. Escape analysis for object-oriented languages. Applications to Java. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 20–34, 1999.

[8] J. Bogda and U. Hölzle. Removing unnecessary synchronization in Java. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 35–46, 1999.

[9] R. Chatterjee, B. G. Ryder, and W. Landi. Relevant context inference. In *Symposium on Principles of Programming Languages*, pages 133–146, 1999.

[10] J. Choi, M. Gupta, M. Serrano, V. Sreedhar, and S. Midkiff. Escape analysis for Java. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 1–19, 1999.

[11] J. Dean, D. Grove, and C. Chambers. Optimizations of object-oriented programs using static class hierarchy analysis. In *European Conference on Object-Oriented Programming*, pages 77–101, 1995.

[12] G. DeFouw, D. Grove, and C. Chambers. Fast interprocedural class analysis. In *Symposium on Principles of Programming Languages*, pages 222–236, 1998.

[13] A. Diwan, J. B. Moss, and K. McKinley. Simple and effective analysis of statically-typed object-oriented programs. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 292–305, 1996.

[14] M. Fähndrich, J. Foster, Z. Su, and A. Aiken. Partial online cycle elimination in inclusion constraint graphs. In *Conference on Programming Language Design and Implementation*, pages 85–96, 1998.

[15] M. Fähndrich, J. Rehof, and M. Das. Scalable context-sensitive flow analysis using instantiation constraints. In *Conference on Programming Language Design and Implementation*, pages 253–263, 2000.

[16] R. Fitzgerald, T. B. Knoblock, E. Ruf, B. Steensgaard, and D. Tarditi. Marmot: An optimizing compiler for Java. *Software: Practice and Experientce*, 30(3):199–232, Mar. 2000.

[17] D. Gay and B. Steensgaard. Fast escape analysis and stack allocation for object-based programs. In *International Conference on Compiler Construction*, LNCS 1781, 2000.

[18] J. Gosling, B. Joy, and G. Steele. *The Java Language Specification*. Addison-Wesley, 1996.

[19] D. Grove, G. DeFouw, J. Dean, and C. Chambers. Call graph construction in object-oriented languages. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 108–124, 1997.

[20] IBM Corporation. *High Performance Compiler for Java*, 1997. `www.alphaWorks.ibm.com/formula`.

[21] D. Liang, M. Pennings, and M. J. Harrold. Extending and evaluating flow-insensitive and context-insensitive points-to analyses for Java. In *Workshop on Program Analysis for Software Tools and Engineering*, pages 73–79, 2001.

[22] T. Lindholm and F. Yellin. *The Java Virtual Machine Specification*. Addison-Wesley, 1999.

[23] J. Palsberg and M. Schwartzbach. Object-oriented type inference. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 146–161, 1991.

[24] J. Plevyak and A. Chien. Precise concrete type inference for object-oriented languages. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 324–340, 1994.

[25] C. Razafimahefa. A study of side-effect analyses for Java. Master's thesis, McGill University, Dec. 1999.

[26] A. Rountev, A. Milanova, and B. G. Ryder. Points-to analysis for Java based on annotated constraints. Technical Report DCS-TR-428, Rutgers University, Nov. 2000.

[27] A. Rountev, A. Milanova, and B. G. Ryder. Points-to analysis for Java using annotated inclusion constraints. Technical Report DCS-TR-417, Rutgers University, July 2000. (Initial report superseded by DCS-TR-428).

[28] E. Ruf. Effective synchronization removal for Java. In *Conference on Programming Language Design and Implementation*, pages 208–218, 2000.

[29] B. Steensgaard. Points-to analysis in almost linear time. In *Symposium on Principles of Programming Languages*, pages 32–41, 1996.

[30] M. Streckenbach and G. Snelting. Points-to for Java: A general framework and an emprirical comparison. Technical report, U. Passau, Sept. 2000.

[31] Z. Su, M. Fähndrich, and A. Aiken. Projection merging: Reducing redundancies in inclusion constraint graphs. In *Symposium on Principles of Programming Languages*, pages 81–95, 2000.

[32] V. Sundaresan, L. Hendren, C. Razafimahefa, R. Vallee-Rai, P. Lam, E. Gagnon, and C. Godin. Practical virtual method call resolution for Java. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 264–280, 2000.

[33] F. Tip, C. Laffra, P. Sweeney, and D. Streeter. Practical experience with an application extractor for Java. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 292–305, 1999.

[34] F. Tip and J. Palsberg. Scalable propagation-based call graph construction algorithms. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 281–293, 2000.

[35] R. Vallée-Rai, E. Gagnon, L. Hendren, P. Lam, P. Pominville, and V. Sundaresan. Optimizing Java bytecode using the Soot framework: Is it feasible? In *International Conference on Compiler Construction*, LNCS 1781, 2000.

[36] J. Whaley and M. Rinard. Compositional pointer and escape analysis for Java programs. In *Conference on Object-Oriented Programming Systems, Languages, and Applications*, pages 187–206, 1999.