A Generating-Extension-Generator for Machine Code

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The problem of “debloating” programs for security and performance purposes has begun to see increased attention. Of particular interest in many environments is debloating commodity off-the-shelf (COTS) software, which is most commonly made available to end users as stripped binaries (i.e., neither source code nor symbol-table/debugging information is available). Toward this end, we created a system, called GenXGen[MC], that specializes stripped binaries.

Many aspects of the debloating problem can be addressed via techniques from the literature on partial evaluation. However, applying such techniques to real-world programs, particularly stripped binaries, involves non-trivial state-management manipulations that have never been addressed in a completely satisfactory manner in previous systems. In particular, a partial evaluator needs to be able to (i) save and restore arbitrary program states, and (ii) determine whether a program state is equal to one that arose earlier. Moreover, to specialize stripped binaries, the system must also be able to handle program states consisting of memory that is undifferentiated beyond the standard coarse division into regions for the stack, the heap, and global data.

This paper presents a new approach to state management in a program specializer. The technique has been incorporated into GenXGen[MC]. Our experiments show that our solution to issue (i) significantly decreases the space required to represent program states, and our solution to issue (ii) drastically improves the time for producing a specialized program (as much as 13,000x speedup).

1 INTRODUCTION

Modern commodity off-the-shelf (COTS) software tends to provide large sets of features to support the diverse use cases of their end-users. However, individual users of many COTS programs might only use a single, fixed subset of the available functionality. From such a user’s perspective, unused functionality constitutes “bloat” in terms of binary size, program performance, and attack surface. A means of producing specialized versions of programs that only include features relevant to a given use case would be a useful tool for simplifying and hardening COTS software. In particular, given certain configuration settings, a developer or administrator may wish to remove features irrelevant to their particular configuration, thereby improving space usage and performance, and reducing the program’s attack surface.

Toward this end, we have created a system, called GenXGen[MC], that specializes stripped binaries. The premise behind our work is that many aspects of the “debloating” problem can be addressed via techniques from the literature on partial evaluation [9, 13]. For instance, a partial evaluator pe takes as inputs (i) a program P (expressed in some language L); (ii) a partition of P’s inputs into two sets, supplied and delayed (for short, S and D, respectively); and (iii) an assignment A(S) to the variables in S. As output, pe produces a residual program P_{A(S)} that is specialized with respect to A(S). More formally, we have

\[ [pe](P, A(S)) = P_{A(S)}, \]

where \([\cdot]\) denotes the meaning function for the language in which pe is written. The requirement on residual program \(P_{A(S)}\) is that it must obey the following equation:

\[ [P]_L(A(S \cup D)) = [P_{A(S)}]_L(A(D)), \]

We find “supplied” and “delayed” to be more suggestive than the standard terms “static” and “dynamic,” respectively.

Here, the partition of P’s inputs is implicit in A(S).

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where \([\cdot]_L\) is the meaning function for \(L\). That is, \(P_{A(S)}\) with input \(A(D)\) produces the same output as \(P\) with input \(A(S \cup D)\); however, \(P_{A(S)}\) has fewer input arguments, and is specialized with respect to the assignment \(A(S)\).

A partial evaluator may be able to identify parts of a program’s control-flow graph (CFG) that are unreachable given particular configuration settings, and produce a residual program that does not contain the identified parts. Moreover, code in the program that is dependent solely on the supplied inputs can be executed by the partial evaluator, and elided from the resulting specialized program. In practice, these abilities allow a partial evaluator to perform a multitude of optimizations, without the developer of the partial evaluator needing to write explicit implementations of each optimization [13]. For example, a partial evaluator will perform removal of unreachable code and constant folding, as well as more sophisticated optimizations, such as loop-unrolling and function in-lining.

For debloating, a partial evaluator can (i) simplify code so that the resulting program incorporates specific features based on particular configuration parameters, and (ii) collapse abstraction layers in the original program via function in-lining.

In some contexts—including in our work—an alternative formulation of the above approach, based on the creation of generating extensions, is more desirable. A generating extension can be thought of as a self-contained, program-specific partial evaluator. A generating extension for \(P\) and a specified set of supplied inputs \(S\) is a program \(ge_{P,S}\) that obeys the following equation:

\[
[ge_{P,S}](A(S)) = P_{A(S)},
\]

where \(P_{A(S)}\) is the specialized residual program defined previously, which obeys Eqn. (2). This approach to program specialization is enabled by a tool called a generating-extension generator: a program that takes as input \(P\) and \(S\), and creates a generating extension \(ge_{P,S}\).

The difference between the two approaches to program specialization can be summarized as follows:

- Applying a partial evaluator to program \(P\) and partial state \(A(S)\) is similar to interpreting \(P\) on an input state, except that the output is a specialized program \(P_{A(S)}\).
- Applying a generating-extension generator to \(P\) is similar to compiling \(P\), except that the outcome is a program, \(ge_{P,S}\), that, when executed on a partial state \(A(S)\), produces the specialized program \(P_{A(S)}\).

Generating extensions, and in particular machine-code generating extensions, have the advantage that they can execute as native programs; a semantic model of the target language is only needed to produce the generating extension. At specialization time, no semantic information is needed (other than the semantics built into the hardware platform on which a generating extension runs).

Moreover, a pre-made generating extension can be delivered to an end user who wishes to specialize a program without needing to deliver additional special-purpose tools for specializing programs. For these reasons, we chose to work with generating extensions.

Program specializers have been created for many different types of languages, including imperative, functional, and logic-based, both for source code and—less frequently—for machine code. However, when creating a program-specialization tool for real-world programs, one faces a multitude of problems. In particular, a program-specialization tool must address two state-management problems:

1. A program specializer needs to be able to save and restore program states efficiently.
2. A program specializer uses a worklist-based algorithm that executes a program over partial program states (§2 and §4). To prevent redundant exploration of the program’s state space, there needs to be an efficient means of determining whether a (partial) state has repeated.

Naive approaches to these issues are extremely costly (§6):
A straightforward approach to issue (1) means copying the entire state for each save and restore operation. The need to test a new state against all states that have previously arisen (issue (2)) suggests the use of hashing. However, resolving collisions requires the ability to compare two states for equality.

These state-management operations have never been addressed in a completely satisfactory manner in prior work, and the disadvantages of prior approaches become more significant in the context of specializing stripped binaries. For instance, programs often use linked data structures, constructed using nodes allocated from the heap. However, for a stripped binary, program states consist of memory that is undifferentiated beyond the coarse division into regions for the stack, the heap, and global data. Moreover, for a program specialist that runs natively, the states that need to be captured and compared in issues (1) and (2) are native hardware states (at the level of the instruction-set architecture).

In this paper, we describe a new technique for state management in a program specialist that runs natively. To demonstrate these techniques, we implement a new program-specialization tool, GenXGen[MC], and present an evaluation of its effectiveness. To address issues (1) and (2), our approach makes use of several ideas known from the literature:

- Using built-in OS process-creation and context-switching mechanisms for saving and restoring states [7].
- Using Rabin fingerprinting to create an incrementally updatable hash of a program’s entire address space where there is an exponentially small probability of the hash of any two states colliding.[19, 21]
- Exploiting hardware support for copy-on-write (CoW) memory management [18] to identify changed memory regions, without the need to instrument arbitrary subject-program memory accesses or resorting to machine-code interpretation. Moreover, the use of CoW reduces physical memory pressure by sharing unchanged pages between multiple processes.

The contributions of our work are as follows:

A. Our main technical contributions are to state management in a program specialist that runs natively (§3 and §5). Unlike prior approaches used in program specializers, our state-management technique does not support a mechanism to resolve hash collisions. Instead, by choosing appropriate values for parameters of the hashing scheme, the probability of a collision can be made arbitrarily small (in our case, $< 2^{-56}$), which allows us to forgo the conventional constraint that collisions be resolvable. By relaxing the collision-resolution constraint to a probabilistic guarantee, we obtain the following benefits:

1. With these technique, state equivalence can be checked in constant time.
2. This state-management technique handles program states over an address space divided into otherwise undifferentiated stack, heap, and global regions. Fine-grained knowledge about variables and types is not required at specialization time. Nor is it necessary for the tool to have knowledge of the distinction between free storage and storage that is in use in the heap.
3. Moreover, we are able to use a hashing technique that supports efficient incremental updating of hash values [6, 22].

B. With these state-management techniques, we implemented GenXGen[MC], a new tool for specializing binaries. We present an evaluation of our technique’s effectiveness in §6. Our approach has several benefits:

- It allowed us to create a program specialist that specializes machine code, runs natively, and can work without symbol-table information (§4 and §5).
- The ability to perform O(1) state comparisons significantly improves specialization performance, compared to a naive approach (§6).
• The use of CoW dramatically reduces memory usage, compared to using full-state copies (§6).

To make the paper self-contained, §2 presents a summary of partial evaluation and generating extensions, using an example to provide intuition. §7 discusses related work. §8 concludes.

2 A PRÉCIS ON PROGRAM SPECIALIZATION

The purpose of this section is to provide background for readers unfamiliar with partial evaluation and generating extensions, and to help them understand how the material in §3–§5 represents an advance over previous work. (Readers already familiar with these techniques may wish to peruse the examples in this section and proceed to §3.) To aid understanding, relevant concepts are presented using source-code examples, using the naive substring-matching procedure match (Fig. 1(a)) as an example.

§2.1 describes how a partial evaluator specializes match on the pattern string. §2.2 describes a C generating extension that performs the same specialization. In both approaches, there is a first phase of binding-time analysis (BTA) and a second worklist-driven specialization phase that produces the residual program. Given the desired partition of the inputs into supplied and delayed sets, BTA extends the partition to the program’s variables at all program points, identifying variable occurrences that can safely be included in partial states. The specialization phase traverses the subject program’s CFG, executing each basic block it encounters. Moreover, the subject program is executed over partial states: states whose values can be safely computed when program execution starts with an assignment to the supplied input variables.

2.1 Overview of Partial Evaluation

The C procedure match in Fig. 1(a) is an implementation of an \(O(|s||p|)\) naive substring-matching algorithm. It returns 1 if and only if the string pointed to by \(s\) contains the string \(p\) as a substring. Note that \(s\) and \(p\) are presumed to point to valid C strings, and thus match terminates whenever the null terminator (ASCII 0) for either string is encountered.

If we partially evaluate match with \(p\) pointing to the string “hat”, we obtain the procedure shown in Fig. 1(b). In this version, the inner loop has been unrolled, and all manipulations and uses of \(p\) and \(s\) have been eliminated: the characters in “hat” are hard-coded into the tests in the specialized

int match(char *p, char *s) {
    while(*s != 0) {
        char *s1 = s; //block 2
        char *pat = p;
        while(1) {
            if(*pat == 0) return 1; //block 3
            if(*pat != *s1) break; //block 4
            pat++; s1++; //block 5
        }
        s++;
    }
    return 0;
}
(a)

int match_s(char *s) {
    while(*s != 0) {
        char *s1 = s;
        if(*s1 == 'h') {
            s1++;
            if(*s1 == 'a') {
                s1++;
                if(*s1 == 't') {
                    return 1;
                }
            }
            break;
        }
    }
    return 0;
}
(b)

Fig. 1. (a) String-matching program match; (b) match partially evaluated on \(p = "hat"\).
procedure. For this example, Eqs. (1) and (2) become
\[ J_{\text{pe}}(\text{match}, [p \mapsto "\text{hat}"] = \text{match}_{p \mapsto "\text{hat}"}, \]
\[ J_{\text{match}}(\text{match}, [p \mapsto "\text{hat}"] \cup A(D)) = J_{\text{match}}(\text{match}_{s} C(A(D)), \]
where \( J_{C} \) denotes the meaning function for \( C \).

Partial evaluation can be implemented using a two-stage process, consisting of BTA and the specialization phase, which specializes the program by executing over partial states [13] (starting with an initial partial state, such as \( [p \mapsto "\text{hat}"] \)).

There are many possible partitions that a BTA algorithm could produce. A BTA algorithm is acceptable for our purposes as long as the partition that it produces for each program point is congruent [13]. Informally, congruence ensures that in every subject-program statement that updates a supplied variable, the update to the supplied variable does not depend on any delayed values. A partition of the variable occurrences at the different program points of \( p \) into supplied and delayed sets (\( V_{s} \) and \( V_{d} \), respectively) is congruent if at every statement \( l \) in \( P \) where a variable \( v \in V_{s} \) is updated, the new value of \( v \) is computed solely from variables in \( V_{s} \). Congruence is important because it ensures that the partial state induced by the set of supplied inputs can always be safely updated.

A BTA algorithm can use forward slicing [12, 28] to compute a congruent partition. Given a set of variables \( V \) and a set of program points \( L \), forward slicing computes the set of program points that may be affected by the values of \( V \) at points in \( L \). For BTA, we compute the forward slice from the delayed inputs. The boxed statements in Fig. 1(a) show the program points included in the forward slice starting at formal parameter \( s \). A congruent partition of the program variable occurrences is implicit in the slice. The forward slice contains all assignments to, and uses of, variable occurrences that are transitively dependent on \( s \), while the complement of the slice contains all assignments to and uses of variable occurrences not dependent on \( s \). Thus, to ensure that the specialization phase only performs safe updates, it executes only the statements in the complement of the slice. Moreover, slicing can be viewed as an extension of BTA results from variable occurrences to statements: all statements dependent only on supplied state are marked as supplied; the remainder are marked as delayed.

The specialization phase is essentially a kind of interpreter that executes \( P \) over partial states, producing a residual program \( P' \). The specializer interprets the CFG of the program, using a partial state to track the values of the variable occurrences in the supplied set. The interpretation is non-standard because at a condition classified as delayed, such as the two boxed conditionals in Fig. 1(a), there are two successor basic blocks to interpret. A worklist is used to keep track of basic blocks that still need to be processed. Every basic block is interpreted linearly, statement-by-statement,
and each statement is evaluated in one of three ways. (1) All statements marked as “supplied” are evaluated, and the partial state is updated accordingly. For example, the statement `pat++` will cause the value of `pat` in the partial state to be incremented by 1. (2) Statements marked as “delayed” are not evaluated, but are emitted to the residual program instead. For instance, the single occurrence of “`s1++`” in the original `match` program is emitted at two different times during the specialization of `match`. (3) However, some statements marked as “delayed” cannot just be emitted as is; if a delayed statement `s` depends on the value of a supplied variable `v`, the value of `v` must be lifted into the residual program’s state at `s`. Lifting can be performed by replacing every occurrence of `v` in the emitted statement with the current value of `v`. For example, lifting is required for the `if` statement in the inner loop of `match`: every emitted instance of the statement in Fig. 1(b) has `*s1` replaced with a character from “hat”.

Unlike a standard interpreter, the specialization phase is prepared to handle control flow governed by delayed state. Consider the `if` statement at the end of the basic block marked as block 4 in Fig. 1(a). Due to the comparison against the (delayed) string pointed to by `s1`, there is not sufficient information in the partial state to determine which branch will be taken. Consequently, the specializer must arrange to specialize the blocks at both successors.

In essence, the specializer needs to “go both ways” when encountering a branch governed by delayed state. In practice, the specializer is generally implemented as a worklist-based algorithm: basic blocks are specialized and residuated using the approach described earlier; however, upon reaching a branch classified as “delayed,” the specializer records the current state, σ, and adds a `(σ, l)` pair to the worklist for every successor block `l`. The specializer then removes an `(s, b)` pair from the worklist, and executes basic block `b`, starting with state `s`. Thus, at the basic-block level, specialization is similar to execution, except that code can also be emitted; at the end of a basic block, the specializer creates the appropriate (partial-state, basic-block) pair(s) for the block’s successor(s), and inserts them into the worklist.

The partial evaluation of `match` illustrates why a partial evaluator needs to be able to check state equality efficiently. Consider block 2, which contains the two assignments at the start of the outer while loop, and ends with an unconditional branch into the inner loop. Every time block 2 is executed, `pat` is set to point to the start of string `p`, and block 3 is enqueued. When block 3 is removed from the worklist, the partial evaluator continues to unroll the inner loop. Subsequently, the partial evaluator reaches the break statement following block 4, triggering a new partial evaluation of block 2: `pat` is reset, and block 3 is again enqueued, ultimately leading to another identical unrolling of the inner loop. Thus, a partial evaluator that always enqueues the successor of block 2, namely block 3, will never terminate.

To prevent this infinite unrolling, the partial evaluator must be able to detect duplicate partial-state/block pairs. In particular, the first time we evaluate block 2, we want to enqueue the pair `(σ, block 3)` consisting of the state `σ` where `p` is equal to `pat` and block 3. Every subsequent time that a partial evaluation of block 2 is complete, we have re-encountered the state-pair `(σ, block 3)`. The partial evaluator will not terminate unless it can determine that `(σ, block 3)` has repeated.

Thus, a worklist-based partial-evaluation algorithm requires two key state-management features:

1. the ability to save and restore partial states,
2. the ability to efficiently check state equality

When partial evaluation is performed on a program written in a type-safe high-level language, both features can be implemented in a relatively straightforward fashion. Assume that `match` is always called such that the pointers `p` and `s` are guaranteed to reference the beginning of valid C strings. In this case, the relevant state is the set of all supplied variables and memory objects reachable from the supplied variables on the stack. States can be saved, restored, and compared
by traversing the graph of memory objects induced by the reachability relation over the supplied state, in a manner similar to the walk performed by a mark-and-sweep garbage collector.

In §5, we describe an alternative method for state management that is more suitable for generating extensions, particularly machine-code generating extensions (§2.2 and §4).

2.2 Overview of Generating Extensions

An alternative approach to program specialization can be implemented via a generating-extension generator. A generating-extension generator GeGen takes as input a program \( P \) and the BTA results, and produces a generating extension:

\[
[\text{GeGen}](P, S) = ge_{P, S}
\]

where the generating extension \( ge_{P, S} \) produces a residual program:

\[
[ge_{P, S}](A(S)) = P_{A(S)}
\]

such that \( P_{A(S)} \) satisfies Eqn. (2). A generating extension has two key advantages over a partial evaluator:

1. It can be implemented as a program that executes natively in the target language, without interpretation. A semantic model of the target language is only needed to construct the generating extension.
2. The structure of a generating extension reflects the basic-block structure of the subject program. Structurally, a generating extension can be thought of as the original subject program, with the partial-evaluation code “compiled in.” This intermingled structure can be structured in such a way that generating extensions can be algorithmically produced basic-block-by-basic-block. Each basic block in the subject program has an associated basic-block procedure in the generating extension that updates the partial state of the subject program and generates residual code. After these actions are completed, the block yields control to the compiled-in state-management logic. This structure was used by Andersen [1] to automatically produce generating extensions for C programs, and we use a similar approach to structure our generating extensions for machine code (but with different state-management mechanisms that are described in §5).

For example, Fig. 2 is an Andersen-style C generating extension for procedure match from Fig. 1(a).

Consider procedure match_ge in Fig. 2: match_ge repeatedly dequeues a (partial-state, basic-block) pair \((\sigma, b)\) from the worklist until the worklist is empty. If \((\sigma, b)\) has not been visited yet, block \(b\) is evaluated on \(\sigma\) through a “basic-block-procedure.”

The state_t struct and snapshot procedure are used to save and restore states. The state_t struct has a member for every variable in the partial state (in this case, just \(\text{pat}\)), and snapshot stores the current values of the variables in the struct. To illustrate the structure of “basic-block-procedures,” we consider three basic blocks from the inner loop of Fig. 1(a): blocks 3 and 4, which end with the two if statements in the inner loop, and block 5, which increments the two string pointers.

The structure of each basic-block-procedure reflects the structure of the basic blocks in the original program. The statements dependent only on the contents of the string pointed to by \(s\) (the boxed statements in Fig. 1(a) and Fig. 2, excluding the if statement) are merely quoted and printed verbatim. Conversely, the statements dependent only on the supplied value \(p\) (the unboxed statements in Fig. 1(a)) are evaluated during the execution of the generating extension. The statement \(\text{if}(*\text{pat} !*= *\text{s1}) \text{break;}\) depends on both \(\text{pat}\) and the delayed string \(*\text{s1}\). Variable \(\text{pat}\) must be lifted: the statement is printed as written, except that supplied variable \(\text{pat}\) is replaced with its current value.

The correspondence between the generating extension and the subject program makes it straightforward to create a generating extension algorithmically. While the execution of the generating
extension is now worklist-driven and non-standard, the basic block-level structure still reflects that of the original subject program. Moreover, the generating extension itself need not include a C interpreter; a special-purpose analysis tool is only necessary to construct the generating extension itself.

Internally, however, the generating extension must still handle the two state-management issues described in §1: (1) saving and restoring partial states, and (2) checking state equality. At the end of each basic-block procedure, the generating extension takes a snapshot of its own state, emits code to transfer control to its successor block(s), and inserts each successor block into the worklist. Finally, control returns to the top of the loop in match_ge, and another state/block pair is dequeued if the worklist is not empty.

3 OVERVIEW

The best prior solution to the state-management problem has been to take advantage of the fact that a partial evaluator is similar to a language interpreter [13]—except that a partial evaluator operates on partial states, and an interpreter operates on full states.

One can design an abstract datatype of partial states for which saving/restoring states and identifying state repetition can be performed with low time and space overhead. In particular, the components of (partial) states can be hash-consed [11] so that a unique representative—i.e., a canonical address—is maintained for each partial state. A set of the addresses of the unique representatives is then maintained, with hashing used to assist membership testing (and collision resolution performed by comparing addresses).

For our work on specializing binaries, such an approach was unsatisfactory. We chose the generating-extension approach, because it creates program specializers that end-users can use without learning sophisticated program-analysis tools. We wanted to avoid (i) packaging a full-featured interpreter for x86 with the specializers or (ii) instrumenting every load and store in the subject binary. Consequently, we did not have the option of implementing memory as an explicit data structure that can be readily swapped to save and restore states—which raises the following question:

How can issues (1) and (2) from §1 be handled efficiently in a generating extension $ge_{P,S}$ that runs natively?

To address issue (1), we use two OS-level mechanisms—copy-on-write (CoW) and process context-switching—to create an efficient mechanism for state-snapshotting and restoration. (See §5.1.) However, the main element that allowed us to devise a solution is that we changed the requirements associated with issue (2) slightly. In particular, we do not insist that there be a mechanism to resolve collisions, as long as we have control over parameters that ensure that the probability of a collision ever arising is below a value of our choosing. (See §5.2.)

In our implementation, by using 128-bit hash values, we ensure that when the program uses $\leq 2^{35}$ bits of memory and visits $\leq 1,000,000$ unique states, the probability of the hash value of any visited state colliding with the hash value of any other visited state is less than $2^{-56}$. Thus, although it is possible for our tool to produce an incorrect residual program due to a hash-value collision, the

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3 More precisely, to support the unique-representative property, one would make use of applicative maps (see [23, §6.3] and [20]), hash-consing, and a hash table to detect duplicates. (The hash-code would be based on the contents of the map’s entries, rather than the structure of the tree that represents the map.)
chances of that happening are negligible.\textsuperscript{4} In our implementation, incremental updating of hash values occurs at page granularity.

Our experimental results show that these mechanisms are critical to the practical tractability of machine-code generating extensions. Most significantly, if Rabin fingerprinting or some other $O(1)$ state-comparison mechanism is not used, the amount of time to produce a residual program scales quadratically with the number of partial states seen. Even on simple examples, fingerprinting yields several orders-of-magnitude of improvement in execution times; one simple test case required over 12 hours without fingerprinting, while requiring only 3 seconds with fingerprinting.

Using CoW produces a large improvement in the amount of space needed to represent partial states; while thousands of pages of memory are required to represent visited states without CoW, at most several hundred are required when CoW is used in our experiments. Moreover, when fingerprinting is used, CoW yields an additional two-to-six-fold reduction in the time required to produce a residual program.

4 CONSTRUCTING MACHINE-CODE GENERATING EXTENSIONS

In this section, we explain the technique for creating generating extensions used in our tool GenXGen[MC]. GenXGen[MC] takes a program and BTA results as input, and transforms each basic block of the program into a self-contained unit that (i) executes the basic block, (ii) updates the partial state that depends on supplied input, (iii) produces a specialized basic block, (iv) snapshots the partial state, and (v) yields control to a controller process. The BTA results determine the transformation of individual instructions performed in step (iii).

GenXGen[MC] relies on the implementation of BTA from WiPER \cite{26}. WiPER invokes CodeSurfer/x86 \cite{3}, which incorporates a number of algorithms for static analysis of machine code \cite{4} to build a dependence graph that supports machine-code slicing \cite{27}. As in §2.1, BTA is performed by slicing forward from the delayed inputs, marking all program points in the slice as delayed, and all points outside the slice as supplied.

To identify lifted values, reaching-definition analysis is performed for the operands of each instruction $I$ in the delayed set. Any static instructions that define an operand used by $I$ must have associated code to lift the value of the operand.

\textsuperscript{4}Put another way, it is as if we had arranged for a year to have the right number of days so that, with any randomly chosen group of 1,000,000 people, the chances of winning a birthday-paradox bet was less than $2^{-56}$.

\begin{verbatim}
  mov dl, [ebx]  -- dereference pat
  cmp dl, 0   -- check first if condition
  jz L7     -- if(*pat == 0) return 1
  mov cl, [eax]  -- dereference s1
  cmp cl, dl  -- check second if condition
  jne L2    -- if(*pat != *s1) break;
  incr eax   -- s1++
  incr ebx   -- pat++
  jmp L3     -- while(!)
\end{verbatim}

Fig. 3. Naive string matcher’s inner loop body. Boxed instructions are delayed, double-boxed instructions have their destination operands lifted, and the remainder are supplied.
In Fig. 3, BTA results are illustrated for the code that implements the innermost loop of match from Fig. 1. Register eax contains the address of the current offset in the string that is being searched, and ebx contains the address of the current offset in the pattern to be matched. The registers cl and dl contain the current characters in the string and pattern, respectively. Basic blocks L3, L4, and L5 correspond to the inner loop blocks in Fig. 1. L7 is reached only if a match is found. L2 is the target of the inner loop’s break statement, which starts another iteration of the outer loop.

In lifted instruction mov dl, [ebx], register dl must be lifted, because cmp cl, dl in block L4 compares the supplied pattern character in dl to a character from the delayed string.

Given the partitioning of instructions in Fig. 3, the generating-extension-generator emits x86 code augmented with pseudo-instructions that expand to sequences of x86 instructions. Their actions emit code, control the flow of computation, and manage partial states.

Fig. 4 illustrates the machine-code generating extension produced from match. In this presentation, we treat the pseudo-instruction actions as black boxes. As before, non-branch instructions classified as supplied are executed, updating the partial state. Conversely non-branch instructions classified as delayed are emitted verbatim in a manner analogous to the printf calls in Fig. 1.

The lifted instruction $I = \text{mov} \ dl, [ebx]$ is executed, just like a supplied instruction. After $I$ is executed, Lift emits code that sets the value of dl in the residual program to the value that dl holds immediately after the execution of $I$.

MakeSnapshot records a snapshot of the current partial state of the subject program, using the technique described in §5.1.

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5The 32-bit Intel x86 instruction set (also called IA32) has six 32-bit general-purpose registers (eax, ebx, ecx, edx, esi, and edi), plus two additional registers: ebp, the frame pointer, and esp, the stack pointer. In Intel assembly syntax, which is used in the examples in this paper, the movement of data is from right to left (e.g., mov eax, ecx sets the value of eax to the value of ecx). Arithmetic and logical instructions are primarily operand instructions (e.g., add eax, ecx performs eax := eax + ecx). An operand in square brackets denotes a dereference (e.g., if a is a local variable stored at offset -16, mov [ebp-16], ecx performs a := ecx). Branching is carried out according to the values of condition codes ("flags") set by an earlier instruction. For instance, to branch to L1 when eax and ebx are equal, one performs cmp eax, ebx, which sets ZF (the zero flag) to 1 if eax = ebx = 0. At a subsequent jump instruction jz L1, control is transferred to L1 if ZF = 1; otherwise, control falls through.

---

L3: EmitSpecLabel(L3)
    mov dl, [ebx]
    Lift(dl)
    cmp dl, 0
    MakeSnapshot
    EmitSpecJmp("jz", L7, L4)
    CondEnqueue("jz", L7, L4)
    Yield

L4: EmitSpecLabel(L4)
    Emit("mov cl, [eax]"
    Emit("cmp cl, dl"
    MakeSnapshot
    EmitDynJmp("jne", L2, L5)
    Enqueue(L2, L5)
    Yield

L5: EmitSpecLabel(L5)
    Emit("incr eax"
    incr ebx
    MakeSnapshot
    Enqueue(L3)
    EmitJmp(L3)
    Yield

Fig. 4. The machine-code generating extension produced for the code in Fig. 3. The three blocks are the machine-code analogs of the handle_block functions in Fig. 2.
Both the Enqueue and CondEnqueue actions place (partial-state, basic-block) pairs into the worklist. In both cases, the state used for every enqueued pair is the state recorded by the most recent call to MakeSnapshot, and a pair is only enqueued if it has not been enqueued before.

Enqueue can be used to enqueue a single successor, in the case of unconditional jumps, or multiple successors, in the case of jumps controlled by delayed state. CondEnqueue is used for conditional jumps governed by supplied state; its action is to enqueue the successor block determined by the supplied state and the condition of the jump instruction.

The actions of EmitJmp, EmitSpecJump, and EmitDynJump are to emit jumps to specialized versions of a block’s successors. If the original jump is conditional and governed by supplied state, we use EmitSpecJump to emit an unconditional jump to the specialized version of whichever successor block is chosen by the jump’s condition. If the original jump is conditional and governed by delayed state, EmitDynJump is used to emit a conditional jump targeting the specialized versions of the block’s two possible successors. If the jump is unconditional, we use EmitJmp to emit an unconditional jump to the specialized version of the only successor.

The action of EmitSpecLabel is to emit a label unique to the current \((σ, b)\) pair being invoked.

The action of Yield is to emit code that yields control to the controller process. The controller process is a simple piece of code that removes a pair \((σ, b)\) from the worklist, restores state \(σ\), and resumes execution at block \(b\).

Pointers. The implementation of the Lift macro needs to correctly handle pointers to stack and heap objects. A pointer to a heap or stack object during specialization time may not be a valid pointer to the object in the residual program. CodeSurfer/x86’s VSA implementation reliably identifies whether a memory location or register holds a pointer to the stack or heap at a given program point in all of the programs tested. By combining this information with a special-purpose implementation of malloc used only in the generating extension, concrete pointers into memory objects can be converted into relocatable offsets.

5 OS-ASSISTED STATE MANAGEMENT

5.1 Implementation of Snapshots

Because a generating extension saves and restores states at the end of every basic block, it is critical for these operations to be implemented efficiently. While saving and restoring CPU state is a straightforward operation, saving and restoring memory state is potentially expensive, both from the standpoint of storage required to represent states, as well as the amount of time required to save and restore states.

We use different OS processes to represent different state snapshots, and the fork() system call to generate a new snapshot of interest. Thus, a set of snapshots—in our case, the elements of the generating extension’s worklist—can be represented by a set of process IDs. A partial state can be restored efficiently merely by performing a process switch.

The fork() system call is implemented in a relatively time and space-efficient fashion through the use of a policy known as copy-on-write (CoW). Through the use of hardware-supported virtual memory, logical addresses used by a process are decoupled from their physical addresses in memory. The address space of a process is broken up into fixed-size pages (typically 4096 bytes), each of which can be mapped to an arbitrary physical memory page. This decoupling lets processes share physical pages, enabling CoW.

When a process \(P\) calls fork(), a new process \(P'\) is created with register and memory contents identical to \(P\), with the exception of eax, which contains the return value of fork(). Rather than allocating new physical pages for \(P'\), every page \(G(P, i)\) in \(P'\) is mapped to the same physical page \(H\) as the corresponding \(G(P, i)\) in \(P\). However, the virtual-to-physical mappings for \(P'\) are flagged as
CoW using hardware support; when $P'$ writes for the first time to a page $G_i(P', i)$ inherited from $P$, a hardware fault occurs, the changed version of the page is allocated its own page $H'$ in physical memory, and the hardware state is updated so that $G_i(P', i)$ is mapped to $H'$.

Consider the two processes $\sigma$ and $\sigma'$ in Fig. 5, where $\sigma'$ is the result of forking $\sigma$ and executing some code that modifies the fourth page of the virtual address space. Each process has a four-page virtual address space, backed by an $n$-page physical memory. In this case, the first three virtual pages of both processes map to the same physical pages, while the fourth page of each process maps to different physical pages.

Our generating extensions exploit CoW to implement the end-of-block worklist update. Each partial state referred to in §2 is a separate Linux process. An additional "controller" process oversees the partial evaluator’s worklist of (partial-state, basic-block) pairs. The controller process serves as a dispatcher for the specialization phase, and the worklist of unprocessed (partial-state, basic-block) pairs is implemented merely as a set of process IDs. Every time the controller process selects a (partial-state, basic-block) pair $(\sigma, b)$ from the worklist, it signals the process $P_\sigma$ that represents $\sigma$, and $P_\sigma$ begins executing. $P_\sigma$ immediately calls $\text{fork}()$, creating a child process $P'$. Initially, the logical address space of $P'$ contains $\sigma$, and its program counter is set to $b$. After $P'$ finishes executing $b$, its logical address space contains $\sigma'$. However, physically only the pages that changed during the execution of $b$ on $\sigma$ are specific to $P'$; the rest are shared with process $P_\sigma$. Because all the memory bookkeeping is implicitly taken care of by the OS and the hardware, the only data state that is stored and manipulated by the controller process is the ID that the OS assigns to each process.

Thus, executing a basic block $b$ on a given state $\sigma$ incurs a cost that is linear in the number of memory-writing instructions in $b$, because each such instruction is executed only once. Moreover, the cost of switching between processes is constant: a switch consists of a system call and the update of several fixed-size hardware registers.

5.2 State Hashing

For a generating extension to avoid traversing previously seen computation paths, it needs an efficient way to determine whether a given state has been seen before. We desire a state-management scheme that possesses five properties:

(i) Given a state $\sigma$ that has been seen before, whenever we encounter $\sigma$ again, the procedure must recognize $\sigma$ as a previously visited state, with no false negatives.

(ii) The procedure must be space- and time-efficient. We would like to store at most several hundred bits of information per state visited. We would also like $O(1)$ state-equality checks.

(iii) The false-positive rate must be kept acceptably low.

(iv) We would like to be able to efficiently update the value that characterizes a visited state.

(v) If the initial state is a multi-gigabyte address space, computing the initial value that characterizes the state should require a constant amount of computation.

These criteria naturally suggest a solution based on hashing. Note that item (iii) deviates from the conventional approach to state management in program specialization, which is that there should be no false positives. As discussed in §1 and explained below, although it is possible for our tool to

Fig. 5. States $\sigma$ and $\sigma'$ after a CoW fault at $P_3$. 

$S_0$ $S_1$ $S_2$ $S_3$ $S'_0$ $S'_1$ $S'_2$ $S'_3$

$P_0$ $P_1$ $P_2$ $P_3$ $P_4$ ... $P_n$
produce an incorrect residual program due to a hash-value collision, the chances of that happening are negligible. Our choice to work with this relaxed requirement was motivated by the fact that our generating extensions work with native hardware states.

To produce efficient generating extensions, item (iv) is especially important. When computing the post-state hash after executing a single block, we wish to perform an amount of computation proportional to the number of changes to the pre-state made during the block’s execution. By exploiting properties of CoW, it is possible to satisfy item (iv) with an appropriate choice of hash algorithm:

- When a basic block \( b \) is executed by the specializer, the first write to a page in the execution of \( b \) induces a CoW fault. Given a log of all CoW faults that occur during the execution of the basic block, we can compute the changes between the pre-state \( \sigma \) and the post-state \( \sigma' \). Implementing this log was straightforward: we added (i) a small amount of instrumentation code to the Linux kernel’s page-fault handler, and (ii) a small amount of extra state to every process structure in the kernel.

- We need an incrementally updatable hashing algorithm—one that lets us efficiently incorporate differing pages into the pre-state hash-code, without additional computation beyond processing of the data of changed pages in the pre- and post-states. Rabin’s fingerprinting scheme\(^6\) satisfies this criterion [6, 22].

Formally, given the pre-state \( \sigma \) and its associated hash \( H(\sigma) \), and the contents of the changed pages, \( P_{\text{pre}} = \{P_1, \ldots P_n\} \), \( P_{\text{post}} = \{P'_1, \ldots P'_n\} \), from the pre- and post- states, we compute the post-state hash using only the pre-state hash \( H \), and the contents of \( P_{\text{pre}} \) and \( P_{\text{post}} \):

\[
H(\sigma'') = H_{\text{incr}}(\sigma, P_1, P'_1, \ldots, P_n, P'_n)
\]

Given a bit-string \( \sigma = (s_0, s_1, \ldots s_{m-1}) \), representing the contents of a program’s address space, we wish to compute a hash \( H(\sigma) \). To do so, the fingerprinting algorithm treats \( \sigma \) as a polynomial \( \sigma(t) = s_0 * + t^1 * + s_1 * t^1 + \ldots + s_{m-1} * t^{m-1} \) of degree \( m - 1 \) with coefficients over \( \mathbb{Z}_2 \). The fingerprinting scheme selects an irreducible polynomial \( P(t) \) (i.e., \( P(t) \) is only divisible by 1 and itself) of degree \( k \), again with coefficients from \( \mathbb{Z}_2 \). Given \( P(t) \), the fingerprint \( H(\sigma) \) is defined as

\[
H(\sigma) = \sigma(t) \mod P(t)
\]

The choice of the degree \( k \) of the irreducible polynomial \( P \) allows us to choose the size of the hash, and thereby tune the collision probability relative to a definition of a “reasonable” execution of a program specializer. For our purposes, we are assuming that the partial state of the subject program will use a 2\(^{35} \)-bit address space and will visit 1 million unique states. Given these assumptions, simple counting arguments outlined in [6] and [22] show that for a 128-bit hash code (i.e., \( k = 127 \)), the probability of there being any collision among the 1 million hash-codes is less than \( 2^{-56} \).

Note that the evaluation of the polynomial at a value of \( t \) plays no part in fingerprinting; we merely use the algebraic properties of the polynomials themselves. Specifically, polynomials with coefficients over \( \mathbb{Z}_2 \) have several properties convenient for the implementation of an incrementally updatable hash:

1. The addition operation \( + \) for such polynomials is addition mod two with no carry—i.e., bitwise exclusive-or. Consequently, subtraction for polynomials with coefficients over \( \mathbb{Z}_2 \) is simply addition. These properties let us treat the contents of memory as \( \sigma(t) \), thus incurring no additional space overhead.
2. Multiplication by \( t^i \) can be implemented as an \( i \)-bit shift.
3. Fingerprinting is linear:

\[
H(A + B) = H(A) + H(B)
\]

\(^6\)Though Rabin fingerprinting is most well-known for its use as a sliding-window hash, it can also be used for incremental hashing [22].
(4) The fingerprint of the product of \( t^i \) and a polynomial \( \sigma(t) \) can be computed via
\[
H(t^i \ast \sigma(t)) = H(H(t^i) \ast H(\sigma(t)))
\]
Given property (1) of polynomials over \( \mathbb{Z}_2 \), in what follows, we will use \( \sigma \) to denote both the bit-string representation of \( \sigma \) and the polynomial \( \sigma(t) \); the intended use will be clear from context.

Consider the pre and post-state in Fig. 5, where virtual page 3 maps to physical page 3 in the pre-state and physical page 4 in the post-state. Properties (1)-(4) admit a simple update procedure:
\[
H(\sigma') = H(\sigma) + H(P_3) + H(P_4)
\]

From (1), it follows that given a change to the \( i^{th} \) page in pre-state \( \sigma \), the post-state \( \sigma' \) can be derived by subtracting off the terms representing the contents of the \( i^{th} \) page in \( \sigma \) and adding on the terms corresponding to the post-state version of the page in \( \sigma' \). By coupling this observation with properties (2), (3), and (4), it can be shown that the \( H(\sigma') \) can be directly computed from \( H(\sigma) \) using only the contents of the \( i^{th} \) page in \( \sigma \) and \( \sigma' \), avoiding the need to examine all of \( \sigma' \) to compute its hash value.

In particular, let \( w \) be the page size in bits supported by the OS (here \( 4096 \ast 8 = 2^{15} \) bits). In addition, let \( \sigma_{a,b} = s_a + s_{a+1}t + \ldots + s_b \ast t^{b-a} \) denote the bit-string containing the bits of the substring of \( \sigma \) starting at \( a \) and ending at \( b \), inclusively, for both \( a \) and \( b \).

Then, from properties (1) and (3), we have
\[
H(\sigma') = H(\sigma) + H(t^{i+w} \ast \sigma_{i+w,(i+1)\ast w-1}) + H(t^{i+w} \ast \sigma'_{i,(i+1)\ast w-1})
\]
and by property (3)
\[
H(t^{i+w} \ast \sigma_{i+w,(i+1)\ast w-1}) = H(H(t^{i+w}) \ast H(\sigma_{i+w,(i+1)\ast w-1}))
\]
For a fixed page size of, e.g., 4096 bits, the only non-constant-time computation is \( H(t^{i+w}) = t^{i+w} \mod P \), which can be computed in time \( \log_2(i+w) \) using modular-exponentiation-via-squaring. Because the maximum amount of addressable memory is bounded on x86 CPUs, \( \log_2(i+w) \) is effectively a small constant in practice.

The number of pages that must be hashed in order to compute the post-state hash is \( O(m) \), where \( m \) is the number of unique pages written during the execution of the basic block. In the common case, \( m \) at most \( O(n) \), where \( n \) is the number of instructions in the basic block. Thus, for the common case, hashing induces a constant overhead on the amount of computation performed by a basic block. The only exceptions are special x86 opcodes, such as those that use the `rep` prefix; these instructions essentially implement loops that perform memory writes repeatedly, until some condition is met. These instructions are often used for, e.g., string operations. In the programs we examined, the use of `rep`-prefixed instructions to write large stretches of memory is uncommon; we did not encounter any cases where `rep` was used to write regions larger than a page. Additionally, in our semantic model we consider `rep`-prefixed instructions to be loops, and we treat the individual prefix-free version of the instruction as a basic block.

In addition to efficient incremental updates, this hash technique also handles new and empty pages efficiently. Any new physical memory added to a program’s address space is zeroed out by the OS for security reasons; thus, when a program’s address space grows, it will contain zeroes. It is clear from the properties of reduction modulo a polynomial that the hash of a zero page is zero; thus, no additional computation needs to be performed to incorporate new pages into the hash of a program state.

6 IMPLEMENTATION AND EXPERIMENTS

Implementation. The work to create GenXGen[MC] was reduced by adopting the BTA implementation from the WiPER partial evaluator [26], which uses the slicing facilities provided by CodeSurfer/x86 [3]. After BTA is performed on a program, GenXGen[MC] traverses the program’s
Generating Extension Generator

CFG, and emits the generating extension using the macros described in §4. The generating extension is then assembled: the generating extension is placed as inlined assembly inside a small C++ wrapper, so the “assembler” is actually g++. The generating-extension binary can then be given values for supplied inputs, and run on a version of the Linux 4.4.14 kernel modified to track CoW faults. The final residual program is then assembled—this time using inlined assembly inside a small C wrapper, so the assembler is gcc. (The use of g++ and gcc for assembly is an implementation expedient.)

**Experimental Questions.** Our experiments were designed to answer the following questions:

1. What are the individual improvements to memory usage contributed by CoW and fingerprinting?
2. What are the individual improvements to the time needed to emit a residual program contributed by CoW and fingerprinting?
3. Compared to the original subject program, how much does specialization speed up execution?

### 6.1 Specialization Performance

**Experimental Setup.** We evaluated GenXGen[MC] using the binaries of seven microbenchmarks—listed in Fig. 6—and, as “real-world” examples, three command-line binaries: two GNU coreutils programs, and one program that makes use of printf. Five of the microbenchmarks were previously used to evaluate WiPER [26]. The sixth, matcher, is the naive string matcher given in Fig. 1. The seventh, stack, is designed to stress test the fingerprinting technique.

- gnu-wc counts lines, chars, or words in stdin. The supplied input specifies which quantities are counted; the delayed input is stdin.
- gnu-env runs a program with a specified assignment to environment variables. The supplied input is the assignment to environment variables; the delayed input is the program to invoke.
- printf is a program that calls into a simple printf library. The supplied input is a format string; the delayed input is the remaining arguments.

These programs present a reasonable cross section of real-world specialization tasks. gnu-wc represents a feature-removal task, in which a single mode of operation is chosen out of a set of potential modes, while printf is a fairly representative layer-collapsing and loop-unrolling task, in which a library call is in-lined into a program. The third program, gnu-env, features aspects of

| Application | Description | Static Input |
|-------------|-------------|--------------|
| power       | Computes \(x^n\) | \(n = 100\) |
| dotproduct  | Computes the dot product of two \(n\)-dimensional vectors | \(n = 100\), and coefficients of first vector |
| interpreter | Interpreter for the minimalist language “Brainf*ck” | an input program |
| filter      | Applies \(m \times m\) convolution filter on an image of size \(n \times n\) | \(m = 3\), \(n = 3\), and elements of the filter |
| sha1        | Computes the sha1 digest of a message of size \(n\) bits | \(n = 1024\) and the contents of the first 512 bits |
| matcher     | A naive substring-matching algorithm | the target substring |
| stack       | A program that writes every stack page \(n\) times | \(n\) |

Fig. 6. Microbenchmarks used in the evaluation.
both tasks, because the core environment-update loop is unrolled, and features corresponding to unused command-line flags are excised.8

For gnu-wc, specializing with respect to the supplied input selects one of three main application loops, each of which is optimized for a different counting task. The generating extension elides the other two loops.

In the case of printf, the specialization unrolls the format string, eliminating run-time parsing and logic for unused format specifiers. Similarly, in gnu-env, the argument-parsing loop is unrolled, emitting a program that runs a program in a pre-defined environment.

To evaluate questions (1) and (2), we implemented GenXGen[MC] so that CoW and state fingerprinting can be independently disabled in generating extensions, yielding four possible execution modes (see Fig. 7).

To simulate disabling of CoW, we added a mechanism to force the copy of an entire process address space. When CoW is “disabled,” we dirty each page without altering the state by (i) writing a single byte to each page in the address space, and then (ii) reverting the page back to its original state. These actions force every page to incur a CoW fault, causing the OS to create a copy of every page in the address space. This approach provides an upper bound on the time required because, by forcing the CoW mechanism to make the copy, a page fault must be handled by the kernel for every page, adding some overhead. We chose to estimate the cost in this way because our generating extensions are inherently multi-process: each process holds a single state. Implementing a true CoW-free approach would have required modifying the OS to eliminate CoW, which seemed unwarranted, given that the technique is not likely to be competitive.

To disable fingerprinting, we implemented an alternative version of the generating extension’s state-comparison and worklist-management algorithm. Without fingerprinting, the only way to compare the states of two processes is to do a direct comparison of process memory. Moreover, we no longer have a convenient means of indexing into a table of previously seen states. Consequently, the state manager must retain a process for every state previously seen, and must compare every newly created process state with every retained state, comparing full address spaces. In contrast, in the fingerprint-based approach, we only need to store the 128-bit fingerprint; any process that does not have outstanding worklist entries can be garbage-collected.

To measure memory usage, the generating extension tracks the number of pages in use across all processes in the generating extension. Because all processes must be retained when fingerprinting is disabled, determining the memory usage across all processes is straightforward: it is the sum of all live pages across all processes. When fingerprinting is used, memory usage is the maximum number of live pages at any given point in the program’s execution. To evaluate the execution time of a generating extension, we time its end-to-end execution, from the beginning of the first basic block to the end of the last basic block.

We allowed the generating extensions to run end-to-end for the “real-world” examples. However, for the microbenchmarks, we added a time-out after 90 minutes of specialization.

Results. The experimental results for questions (1) and (2) are presented in Fig. 7. With respect to question (1), both fingerprinting and CoW play a significant role in reducing memory usage. Using CoW, however, yields the most significant reduction for every application, except stack. This improvement is due to the fact that for all ten applications, the instructions that are evaluated

8The reason we used only three real-world programs in our study was because of limitations of CodeSurfer/x86 [3, 4], which GenXGen[MC] uses to implement BTA. Fortunately, in many circumstances, the subject program can be adapted to overcome the limitations, e.g., by manually unrolling a loop. However, the effort required to identify appropriate rewritings to overcome current limitations of the static analyses in CodeSurfer/x86, as well as to model calls to library functions, limited the number of real-world programs that we were able to use for our study.
## Generating Extension Generator

### Generating-extension performance

| [CoW,Fingerprint] | orig. | resid. |
|-------------------|-------|--------|
| [no,no]           |       |        |
| [yes,no]          |       |        |
| [no,yes]          |       |        |
| [yes,yes]         |       |        |

### Execution time

| program     | time  | pages |      |      |
|-------------|-------|-------|------|------|
| printf      | 68m   | 240577| 48   | 12774|
| gnu-wc      | 45m   | 146901| 46   | 2129 |
| gnu-env     | 13h   | 958050| 129  | 2129 |
| power       | 74m   | 221416| 102  | 2129 |
| dotprod.    | >90m  |       | 2129 | 1    |
| interp.     | >90m  |       | 2129 | 1    |
| filter      | >90m  |       | 2129 | 1    |
| sha1        | >90m  |       | 2129 | 1    |
| stack       |       | 195900| 1959 | 1959 |

### Fig. 7. Run times and space usage for each generating extension, with and without CoW/fingerprinting. Run times for original and residual programs are also included, with 95% confidence intervals (“—” means “not measured.”)

### Instruction counts

| program     | original | residual |
|-------------|----------|----------|
| printf      | 754      | 1038     |
| gnu-wc      | 1929     | 775      |
| gnu-env     | 1820     | 1123     |
| power       | 30       | 323      |
| dotprod.    | 307      | 1123     |
| interp.     | 146      | 558      |
| filter      | 287      | 1207     |
| sha1        | 332      | 2823     |
| matcher     | 34       | 410      |
| stack       | 3930     | 1        |

### Fig. 8. Instruction counts for original and residual programs.
during generating-extension execution perform the majority of their writes within a single stack page. Even when fingerprinting is not used, CoW ensures that the number of pages needed to retain all previously visited states is small, roughly the number of basic blocks that executed at least one memory write.

Regarding question (2), fingerprinting plays the most significant role in reducing execution time. This result is unsurprising, because the amount of time needed to identify whether a state has been previously visited without using fingerprinting scales linearly with the number of states previously visited. Thus, the execution time scales quadratically with the number of states.

Stress-test stack$(n)$ performs a set of writes that causes the generating extension to recompute each stack-page fingerprint $n$ times. Still, the benefits of $O(1)$ lookup outweigh the cost of repeatedly fingerprinting every stack page.

Using CoW also improves the execution times of generating extensions; the improvement is most pronounced in the case where fingerprinting is also used. When fingerprinting is used, the overhead of copying an entire process begins to dominate the execution time of the generating extension.

For the gnu-wc and stack generating extensions without fingerprinting, the execution time with CoW enabled was greater than when CoW was disabled. We do not have a full explanation, but we believe that the extra cost is due to the cost of collecting memory-usage data. When we measure memory usage with CoW enabled, we track every process currently using a given page. For certain workloads, especially when fingerprinting is not used—and thus page mappings are retained for every state visited—the cost of maintaining this data structure may become relatively large.

### 6.2 Speedup of Specialized Programs

To evaluate experimental question (3), we timed the end-to-end execution time of each program on an input, collecting the 10% trimmed mean of 1001 executions: i.e., for the original and residual version of each program, we ran the program 1001 times, and discarded the 100 shortest and 100 longest execution times.

To time the programs, we instrumented the beginning and end of main in each program with calls to a rdtscp-based timer. By doing so we avoid recording the noise induced by the initial context-switching, loading, and execution of the program. The hardware-counter-based rdtscp counter provides 40-clock-cycle resolution.

For question (3), results are presented in Fig. 7. Specialization produced a speedup in all but one program. Because stack has no meaningful delayed actions, it has a 5500x speedup, due to the elision of thousands of memory writes. The most significant speedup in a specialized program with non-trivial delayed functionality was for matcher, which was 9x faster. This improvement can be attributed to the specialization of the inner loop of the program, which elides all the memory loads for the target string. In particular, this change speeds up the common case in which a character in the string being searched does not match the first character of the target string. filter yields the second most significant speedup, being 6.8x faster. The specialization of filter significantly optimizes the inner loop of the image filtering procedure, eliminating the if statement that selects which image filter is applied to each pixel, as well as inlining loads from lookup tables that encode properties of the selected filter algorithm.

power and dotproduct benefit significantly from the unrolling of their main loops; the elimination of the branch condition at the loop head yield a 5.5x speedup and a 4.6x speedup, respectively.

The specialized version of gnu-wc has a speedup of 2.7x The specialization of gnu-wc elides the argument-parsing loop, as well as setup code that (i) sets locale information and (ii) obtains system-dependent configuration information.
gnu-env, printf, enjoy more modest speedups, roughly 16-18%. Most of the speedup is due to the unrolling of the core loop in each program: for format-string parsing in printf, and the argument-parsing and environment-setup loops in gnu-env.

sha1 obtains a 1.4x speedup from the elision of loads inside the main loop, along with the elision of the initial code that initializes the supplied data.

However, interpreter experiences a slight slowdown (< 1.1x), possibly due to the effects of aggressive unrolling on cache performance.

7 RELATED WORK

Specialization of Machine Code. Run-time code generation is a generating-extension-like approach to program specialization that produces machine code on-the-fly during program execution. Unlike our approach to machine-code specialization, which operates on stripped binaries without source code or symbol-table information, run-time code generation systems take user-annotated source code as input and perform BTA and generating-extension construction as part of compilation. In the Fox [15, 16] and Lancet [24] systems, type-level information in the source code is exploited to produce run-time machine-code generators. These systems avoid the state-management issues from §5 by exploiting the availability of high-level semantic information from the source language. In ’C [8], the user implements code generators using a DSL, and the user has the burden for avoiding redundant states and ensuring that code generators terminate. In contrast, Klimov [14] describes a run-time code generator for Java bytecode that does not rely on information from source code. However, Klimov can only determine state equality for programs that do not use the heap; the approach identifies semantically identical states based on structural properties of Java Virtual Machine heap configurations. JIT compilation [2] is an example of run-time code generation in widespread use. However, because it is performed at run-time, the emphasis is on recouping the cost of translation, which limits the kinds of optimization techniques that can be performed.

Turning to interpretation-based approaches, WiPER [26] and TRIMMER [25] are partial evaluators for x86 binaries and LLVM IR, respectively. WiPER uses CodeSurfer/x86’s semantic models of the 32-bit x86 instruction set to evaluate instructions. WiPER represents states using an applicative-map-based data structure that does not use hash-consing. Thus, state equality is determined by directly comparing the contents of the data structure. TRIMMER implements a non-traditional approach to partial evaluation. In particular, TRIMMER implements an aggressive extension of LLVM’s loop-unrolling and constant-propagation passes, rather than full partial evaluation. By doing so, TRIMMER avoids the need for a general-purpose state-management strategy.

Although our system is quite different from Fox [15, 16] in most respects, their use of pseudo-instruction macros inspired our approach to constructing machine-code generating extensions. We use similar macros to produce residual assembly code, and extended the approach to include various other state-management actions.

Incremental State Hashing. To the best of our knowledge, our work is the first application of incremental state hashing to program specialization. Our fork-based method for managing partial states was inspired by the state-management mechanism in the EXE symbolic-execution system [7].

Model checking is a method to check properties of programs statically by exploring the state space of a transition system. To achieve acceptable performance, model-checking algorithms must avoid exploring redundant states, and Rabin’s fingerprinting technique has been used to implement incremental hashing of program states in model checkers [19, 21]. One of these model checkers, StEAM [17, 19], harnesses a VM that interprets assembly language. In contrast, our implementation
of incremental state hashing exploits OS-level information to apply the technique to code that executes natively, rather than in an interpreter.

**Symbolic and Concolic Execution** Partial evaluation bears some resemblance to symbolic execution. In both cases, the state space of the program is partitioned: into supplied and delayed variables in partial evaluation, into symbolic and concrete variables in symbolic execution. Moreover, we use the OS-based state-management techniques previously implemented in systems such as EXE.

However, symbolic execution differs significantly from partial evaluation in terms of how the partitioned state space is explored. In both symbolic execution and partial evaluation, part of the state space is kept concrete. In the case of symbolic execution, the non-concrete part of the state space is represented as sets of symbolic values, which are logical formulas in some theory. Partial evaluators, on the other hand, do not track any information about the non-concrete part of the state space; the congruence property of the BTA algorithm ensures that the delayed state is never needed to update supplied values.

Symbolic execution thus attempts to construct a symbolic approximation of the values in the symbolic portion of the state at every program point. That is, it attempts to explore, as exhaustively as possible, precisely the subset of the state space that a program specializer ignores. The state-management techniques in symbolic execution are geared towards managing the symbolic state. In particular, we are not aware of symbolic-execution engines that enforce any sort of termination or state-equality properties with respect to the concrete portion of the state. To perform partial evaluation with a symbolic-execution engine, one would need to be able to determine when, for every program point, all concrete states reachable from the starting concrete state had been reached. There is no straightforward way to achieve this with a stock symbolic-execution tool.

8 CONCLUSIONS AND FUTURE WORK

This paper describes how (i) the desire to perform specialization of machine code, using generating extensions running natively, motivated (ii) the development of new techniques for state management in a program specializer. The main challenge was that machine-code programs perform arbitrary reads and writes to an undifferentiated address space, and for this reason, our solution—in part—makes use of existing OS-level functionality. Our technique is used in the generating extensions created by GenXGen[MC].

It has not escaped our attention that our technique can also be used for source-code specialization. In fact, we have already adapted the implementation to create GenXGen[C], a prototype generating-extension generator for C programs.

The state-management technique presented in this paper is not the only option for a source-code specializer: because sufficient information about a program's variables is available, an interpreted approach to source-code generating extensions could track the state of memory at the level of individual variables. However, for source-code specialization, our technique offers three advantages:

- A generating extension for a source-code program is compiled to machine code, and hence specialization is performed by compiled—rather than interpreted—code.
- By intercepting CoW faults and incorporating changed pages into the hash-value for a memory state, a program specializer can easily support programs that use linked data structures, with no need to perform a mark-and-sweep traversal to capture program state. Thus, our technique provides a method that can be used for languages that are not memory-safe, such as C (although a conservative mark-and-sweep algorithm [5] could be employed).
- The state-management implementation can be shared among program specializers for different languages.

As future work, we plan to develop GenXGen[C] further, and to investigate its applications.
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