On-the-fly Code Activation for Attack Surface Reduction

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Abstract

Modern code reuse attacks are taking full advantage of bloated software. Attackers piece together short sequences of instructions in otherwise benign code to carry out malicious actions. Eliminating these reusable code snippets, known as gadgets, has become one of the prime concerns of attack surface reduction. The aim is to break these chains of gadgets, thereby making such code reuse attacks impossible or substantially less common. Previous work on attack surface reduction has typically tried to eliminate such attacks by subsetting the application, e.g. via user-specified inputs, configurations, or features, or by focusing on third-party libraries to achieve high gadget reductions with minimal interference to the application.

In this work we present a general, whole-program attack surface reduction technique called OCA that significantly reduces gadgets and has minor performance degradation. OCA requires no user inputs and leaves all features intact. OCA identifies specific program points and through analysis determines key function sets to enable/disable at runtime. The runtime system, thus, controls the set of enabled functions during execution, thereby significantly reducing the set of active gadget chains dramatically. On SPEC CPU 2017, our framework achieves 73.2% total gadget reduction with only 4% average slowdown. On 10 GNU coreutils applications, it achieves 87.2% reduction. On the nginx server it achieves 80.3% reduction with 2% slowdown. We also provide a gadget chain-breaking study across all applications, and show that our framework breaks the shell-spawning chain in all cases.

1 Introduction

Attack surface reduction has gained in importance lately. A recent investigation [51] showed that on average, 95% of GNU libc code [23] is never used by user applications in a typical Ubuntu Desktop installation. This bloated software has security implications. The excess code can contain bugs and is typically not maintained well. This allows it to become a landmine of vulnerabilities, or to be repurposed for malicious ends in a code reuse attack.

In this work we aim to reduce the exposed, reusable portions of code that are available when an attacker launches a code reuse attack. We motivate our work with a simple but powerful example. It shows why attack surface reduction is important, and also why it is challenging to make a meaningful dent in the reusable code used for staging the attack.

Listing 1 shows an example of an “ROP gadget chain.” These chains can be used as part of an attack to leak secrets, hijack processes, or otherwise cause damage. We provide more background on return-oriented programming (ROP) and gadgets in the next section, but for the unfamiliar, the high level idea is this: Snippets of code (gadgets) are strung together with return statements to jump from one to the next, in order to carry out some malicious computation to launch an attack.

In the above example, the chain is designed to compute
relatively little but to a powerful end: It will launch a shell via `execve`. This is a real chain that has been automatically generated by the Ropper tool [54] on the `nginx` server application [43]. Launching such an attack requires exploiting a memory vulnerability, which is also a realistic possibility here. CVE-2013-2028, for example, is a bug in the decoding functionality of `nginx`. Carlini et al [6] show how to exploit this bug to write to arbitrary locations. Note that as with any code reuse attack that we consider in this work, we assume the attacker has some way of initiating the attack. Here we assume, for example, that the attacker can exploit some memory vulnerability to write this chain into the stack, overwrite the return address to point to gadget 1’s address, and freely return along the ROP chain. Due to C’s weak memory model and high costs of full memory protection, such vulnerabilities unfortunately do exist within almost all software written in C.

Referring to Listing 1, the ROP gadget chain proceeds as follows. It stores the address of the “/bin/sh” string into some location in the .bss section (gadgets 1-3). Then, 8 bytes beyond that, it stores the value 0 (gadgets 4-6). Then it places the address of “/bin/sh” into the rdi register, and it places the 0 value into rsi and rdx (gadgets 7-9). Lastly, it stores the `execve` syscall number 0x3b into the rax register and then executes the syscall instruction (gadgets 10-11).

The actual control flow for executing this attack is return-based. The gadget addresses are the return targets. Thus, each `ret` instruction does the attacker’s bidding, popping the gadget address and jumping to the next malicious snippet.

There are several important details about the strength of gadget chains. First, as in this case, the chain does not need to be Turing-complete to be effective. Here the gadgets perform a specific set of actions, all in the service of setting up an `execve` call that will launch `/bin/sh`. Second, the gadget chain does not need to be built from “intended” instructions on architectures such as x86. Though not shown in the listing, gadget addresses from this `nginx` example leverage “unintended” instructions. They are at unintended offsets in the original instructions (which is allowed in x86’s variable-length instruction set architecture). Third, the gadgets are not particularly rare instruction sequences, so scavenging these gadgets (or computationally equivalent sequences) from the executable code ought to be achievable given a sufficient set of instructions.

These points illustrate the importance as well as the difficulty of the problem. Not only is it possible to create an attack from only a handful of gadgets from a large pool of instructions, but on x86 the gadgets can be cherry-picked from what is essentially a byte stream of executable code. And this attack represents a very serious security breach. Once the attacker hijacks the process and launches a shell, they can perform any action with that process’ permissions.

Defending against code reuse attacks is still an open research problem. Though we have made several assumptions in this simple example, more complex gadget-based attacks exist. Current defenses are not able to handle them. A critical question is whether the attack surface can be reduced substantially enough to break such chains.

Thus, we propose OCA, an attack surface reduction technique for reducing reusable gadgets and breaking their chains. This paper makes the following contributions (please refer to Section 5 for results and their analysis):

1. The first sound, whole-application technique for on-demand loading and purging for attack surface reduction which also has low runtime overhead.

2. An evaluation on SPEC CPU 2017, GNU coreutils, and `nginx` that is comparable to unsound techniques in terms of gadget reduction.

3. Evidence that short but highly detrimental gadget chains can be broken by this technique.

In the next section we present details of the problem and closely related solutions. Then we give an overview of our solution in Section 3, including assumptions. Then we provide details of the framework in Section 4, followed by an evaluation in Section 5. Lastly we provide more related work (Section 6) and conclude (Section 7).

2 Background and Motivation

Code injection could be considered the predecessor of modern code reuse attacks. Early code injection attacks could simply inject code into memory such as the heap and execute it. This was countered by data execution prevention (DEP) [2]. DEP enforces the write XOR execute (W ⊕ X) property on pages, which is sufficient for stopping such blatant attacks.

Attackers grew to overcome this. Perhaps the most basic code reuse attack is the classic return-to-libc attack [42, 59]. In a return-to-libc attack, the attacker exploits some memory vulnerability to compromise “control data,” i.e. an indirect branch instruction (return, indirect jump, or indirect call). Then the attacker redirects control flow to jump into GNU libc (glibc) code, which is full of functionality that can be very useful to an attacker. For example, one useful place to jump can be the `mprotect` function. If the attacker can control the parameters to this function, then they can remap a page as writable, breaking the W ⊕ X property, and allowing them to carry out code injection. Address space layout randomization (ASLR) [47] can make it more difficult to locate target code such as the glibc library calls, but then there are also known attacks that get around it [17, 20, 56, 58].

Reuse attacks have only become more complicated. Return-oriented programming [55], jump-oriented programming [5, 8], and call-oriented programming [53] are all techniques that leverage existing code to perform attacks. They rely on “gadgets” in the code base, which are sequences of instructions that can be strung together to perform some computation. In
fact, an attacker can construct a Turing-complete program from gadgets that exist in the code base, but this is often not even needed. Turing-incomplete functionality may be sufficient if it still allows the attacker to carry out some useful action (e.g. leaking a secret, which could be the end goal, or spawning a shell, to give the attacker more control).

Traditional defenses to these attacks have had some success, but they have also had their shortcomings (see Section 6). Due to these and other reasons, attack surface reduction is one class of defense that has gained in prominence lately. Piece-wise compiler [51], Chisel [24], Razor [49], and BlankIt [48] are four recently developed debloating/attack surface reduction techniques that motivate our work. They successfully show how to reduce applications’ attack surfaces, but they have shortcomings, as well.

Piece-wise compiler [51] modifies the loading stage at process start-up to remove unreachable library code. It is sound, removing only unneeded functionality from libraries (for which it requires libraries’ source code). No user input is needed, but piece-wise-compiled libraries must be provided to a program before running it. This approach removes function(s) in the library only if they are proved unreachable on a whole-application basis, i.e. from nowhere in the driving application can they ever be invoked. Due to complex control flow in the libraries and conservative limitations of static analysis, proving such a property becomes extremely difficult. As a result this work has not been demonstrated on glibc, a real world library with vulnerabilities; on musl-libc its success is limited, leading to debloating roughly 73% of its gadgets when linked with SPEC CPU 2006.

Chisel uses reinforcement learning to learn which parts of a program are actually used and needed, and then builds a trimmed version of it. Chisel is an unsound technique. It may learn a model that eliminates needed functionality. This can induce crashes and so is not practical in the general case. Chisel works in a kind of compile-test-refine loop, as its learner identifies which parts of the program are needed in order not to crash or provide bad output. As such, it requires a user specification (supplied as test inputs) and source code. Chisel is designed to work with application code.

Razor uses heuristics and test inputs to debloat binaries. Like Chisel, Razor is unsound and can lead to crashes. Razor is designed for and works on applications, though they include a discussion on its current effectiveness on libraries. Similar to Chisel and Piece-wise, it also cannot debloat “may-use code.” That is, when these techniques remove a piece of code, it is because that code is presumed unnecessary.

BlankIt is a binary runtime technique for libraries that loads the library code that is needed on each call site in a dynamic sense, i.e. on each call to a library from the driving application. BlankIt achieves a very high precision of attack surface reduction of 97.5% and removes all known CVEs from real world libraries such as glibc. It is able to achieve this precision due to call-specific predictions of reachability that stem from the ML-based analysis of its arguments. It is able to thus carry out debloating of “may-use code” due to runtime analysis. It accomplishes attack surface reduction by using a binary runtime framework to load and unload library code as needed. The runtime also has a predictive component that can flag unexpected control flow. BlankIt is however limited only to libraries and does not work on application code.

We summarize and compare the characteristics of these four approaches in Table 1. With regard to practical usage, we find the last two rows strikingly important (i.e. whether it is sound, and whether it can debloat may-use code). Regarding soundness, this is critical for a debloating technique to achieve general adoption. It cannot make unsound transformations that can crash applications. Regarding whether the technique can debloat may-use code, this is critical for diminishing the “blast radius” of an attack. Solutions such as Razor, Chisel, and Piece-wise do not reduce the attack surface for any code that may be needed.

Table 1: Broad properties of 4 state-of-the-art debloating techniques for security (Piece-wise, Chisel, Razor, and BlankIt).

|                              | PW | Chsl | Rzr | BI |
|------------------------------|----|------|-----|----|
| Works on application        | ✔  | ✔    | ✔   | ✔  |
| Works on library            | ✔  |       | ✔   | ✔  |
| Works on binary             | ✔  | ✔    | ✔   | ✔  |
| No user input needed        | ✔  | ✔    | ✔   | ✔  |
| No training needed          | ✔  | ✔    | ✔   | ✔  |
| Is sound                    | ✔  | ✔    | ✔   | ✔  |
| Can debloat may-use code    | ✔  | ✔    | ✔   | ✔  |

Some of the above problems stem from the underlying techniques used, especially involving machine learning. One of ML’s typical drawbacks is identifying realistic and plentiful training data. This problem is present in both BlankIt and Chisel. Another drawback can be the negative impact of mispredictions. In Chisel, the prediction is at compile-time, and misprediction leads to soundness issues. In BlankIt, prediction is at runtime, which has two issues: (1) how to act when an alarm is raised (which could be a false positive); and (2) how to ascertain whether the input data to the predictor is trustworthy. The inputs to predictive models are susceptible to memory corruption vulnerabilities, just as other parts of the program are. BlankIt attempts to alleviate the problem by hoisting the values and their predictions as high up in the control flow as possible which can potentially help with this, but it is still not a guarantee.

Both Chisel and Razor require a user specification of what functionality an application needs. In both cases this specification is provided as test cases. Not only can this requirement limit adoption for real-world use, in both of these works it is tied directly to the soundness. If the user does not provide the “right” set of test cases, then the debloated program may
crash. In their evaluation, several such cases are shown which is a significant issue.

To the best of our knowledge, there is no general technique today that (1) works on the applications as a whole instead of libraries, (2) is sound, and (3) can debloat may-use code. Current techniques either tackle libraries to achieve strong attack surface reduction, or they tackle applications and compromise soundness. Furthermore, an ideal solution would not require any test cases or specification from the user; and it would either avoid prediction or handle its security challenges gracefully. All these limitations motivate the current work.

3 Overview

3.1 Proposed solution

We propose a compiler and runtime solution for attack surface reduction called OCA. It is sound, has a static and runtime component, requires no user input, requires no hardware changes, and works on application code.

OCA embraces the idea that only code that is currently needed by a running program should be available for execution; the rest should be made inaccessible such that any attempt to access it should trigger a runtime exception. Active sections of code form “decks” that the program can effectively stand on. When a deck is unneeded, it can be removed. To take the analogy further, OCA is a technology for attack surface reduction, but it can be viewed as constructive. It achieves attack surface reduction not by cutting down the program, but by putting together the active code that it needs at any particular execution point. A deck could be a group of functions which are guaranteed to be executed from current execution point. Since such a set cannot be precisely generated without causing heavy runtime overheads (especially inside loops), OCA will turn this into a tight overestimation problem inside respective regions.

OCA depends heavily on static analysis. The decks of the program are based off of static features. An outermost loop in a loop nest must be treated as its own deck, for example. Similarly, a single-function leaf node in the callgraph, which is not reachable by any loop, forms its own simple deck. In our implementation, granularity of disabling/enabling mechanism is at the system page level. Creating and tearing down a deck corresponds to marking code pages read-execute (RX) and read-only (RO), respectively. In other words, if a deck consists of functions foo() and boo() is to be turned on or enabled at some program point, one must execute calls to mark the respective pages that contain foo() and boo()’s code as read-execute (RX).

Figure 1 shows a high-level view of the compiler step of the OCA solution. Application source code is fed into the LLVM compiler pass. The compiler performs static analysis and identifies programs points for decks to be enabled (RX) and disabled (RO); it also performs partitioning of such sets of functions to improve security benefits as discussed later. Additionally, the pass creates a custom linker script. Both of these are fed into the linker (where the -T option consumes the custom linker script). The linker produces the final binary. OCA is represented by the blue part in the figure, and the linker itself is unmodified.

Figure 2 illustrates the runtime by way of an example. It is drawn directly from the GNU coreutils’ date program [10], which provides a command-line option for reading dates from a file (given by the -f switch). When this option is given, main invokes batch_convert. At runtime, OCA will create and tear down a deck for this single function, batch_convert. In Figure 2, there are two code pages in memory (for simplicity). Page A contains main, and page B contains batch_convert. The call to batch_convert has been instrumented by the compiler, so that before it is invoked, its page will be mapped RX, and after it returns, its page will be mapped RO. These mapping steps are done by deck_single and deck_single_end, respectively.

bc_funcid is the function ID for batch_convert assigned by OCA at compile-time. The 4 program points, P1-P4, indicate which pages are mapped RX at each step. One can see that gadgets in unmapped decks and their pages are thus inaccessible to the attacker; thus, the finer the granularity of the deck (finest granularity being one function), the better the security.
3.2 Threat Model

We assume the operating system and compiler are trusted. The source code and any third-party libraries may contain bugs. For simplicity, we do not handle dynamically generated code or self-modifying programs; we focus on C/++ binaries.

We are focused solely on attack surface reduction and assume the attacker has some way of initiating and propagating the attack (e.g., that the attacker can exploit a memory vulnerability and trigger a gadget chain). Given today’s state-of-the-art defenses, we find this assumption reasonable.

We assume the runtime is protected (a similar assumption in [48]), and which can be implemented with in-process isolation [30], hardware segmentation or software fault isolation [31]. This prevents attackers from jumping into the runtime and guarantees, along with the trusted loader, that the statically computed function IDs and framework metadata are protected.

Arguments to the runtime API are statically evaluated and passed by register and so cannot be tampered with, except in the case of indirect calls. How to guarantee the integrity of indirect call targets is precisely the problem handled by orthogonal schemes like CFI and CPI (see Section 6). OCA does not tackle this problem, which we consider out of scope. OCA is focused on reducing the attack surface available to an attacker when an attack occurs, but schemes such as CPI would still be needed for pointer integrity. Similarly, redirecting control flow to another instrumented runtime call that is mapped RX depends on orthogonal defenses. Repeatedly invoking the same instrumented runtime call that is mapped RX, however, is disallowed by construction (see Section 4, which details how instrumented calls will only execute exactly once for every paired teardown call).

As described earlier, the threat is an attacker exploiting the memory vulnerabilities of an application executing under the OCA system, attempting to string together a gadget chain to launch an attack. Due to needed gadgets residing in multiple decks that are disabled, however, the attack will lead to a runtime exception and be caught.

4 Framework

The framework consists of a compiler and runtime component. The compiler is responsible for inserting calls in the application to the runtime in order to create and tear down decks. The runtime receives requests from the application and enables and disables the code pages associated with each deck. In this section we discuss each part separately and then cover optimizations.

4.1 Compiler component

The compiler portion is an LLVM [32] pass that can be divided further into two parts: instrumentation and linker script output. During instrumentation, OCA identifies function calls and loops and instruments them appropriately with calls to the runtime. As the pass does this, it collects critical static information for organizing the text section, which it uses to create a custom linker script. The key idea of this work is to keep the decks\(^1\) as lean as possible. Ideally, each deck should consist of one function, i.e., only one function should be enabled at a time for highest security. Such a scheme would incur very high overheads, however, especially for call chains that execute inside loops, and would make the scheme untenable; thus, based on the context surrounding a program point, static analysis identifies what a deck should be and hoists the calls to the runtime accordingly.

4.1.1 Analyzing for Decks

Analysis and instrumentation for decks is heavily organized around loops. Loops are problematic because adding code for enabling or disabling a deck inside them can cause significant performance degradation. We define two terms: encompassed and non-encompassed functions to distinguish between the loop context that surrounds them. A function is encompassed if it is called inside of a loop, or if it is reachable within the callgraph by some function that is called within a loop through a caller-callee relation. To determine the encompassed function set, the pass first identifies all functions called within a loop, and then takes the transitive closure of any functions reachable from that set using the caller-callee relation shown by the callgraph. The non-encompassed function set is simply the set of all functions minus the encompassed function set.

OCA’s default treatment of loops is to bear on the side of performance. Due to this reason, it tries to avoid instrumenting inside of them, because if this function is called in a surrounding inter-procedural loop at runtime, its deck’s instrumentation will incur repeated invocations, leading to high overheads. Interprocedurally this implies that it cannot instrument inside of encompassed functions, either. This also raises a problem for loop-enclosed indirect calls, whose static target set can be large, and whose precise dynamic value is often unknown until execution is inside of the loop.

To handle these different cases, the pass instruments four different types of decks which will invoke the runtime: (1) Single, (2) Loop, (3) Reachable, and (4) Indirect. We describe each one, covering where it is placed, and what it is responsible for mapping.

The single deck is used when a non-encompassed function calls a non-encompassed function. This is the simplest case. The compiler must ensure that the function being called is mapped RX before it is actually executed. Because the callee is known to be non-encompassed (i.e., not part of some transitive closure that lies within a loop), only a single function

\(^1\)A deck is defined as a group of functions that are enabled at a program point by turning their page permissions to RX.
needs to be mapped RX (i.e. the callee itself).

The loop deck is placed at the outermost loop header for any loop nest in any non-encompassed function. It is designed to map all functions that can be reached interprocedurally within that loop. Notice that an encompassed function will never instrument a loop deck. If it did, it could be invoked repeatedly within a possible interprocedural loop, which could lead to severe slowdown.

The reachable deck is placed in a non-encompassed function before any calls to encompassed functions. Even though encompassed functions are, by definition, part of some interprocedural loop, they may also be reachable in the callgraph via some non-loop path. Encompassed functions should also not contain any instrumentation (due to the potential performance degradation). Thus, when a non-encompassed function invokes an encompassed one, the compiler needs the reachable deck to map the encompassed function and any interprocedurally reachable functions from it as RX.

The indirect deck is for function pointers. The challenge of function pointers is that their exact targets are often not known statically, and therefore the compiler cannot determine precisely what needs to be mapped RX until runtime. Function pointer analysis can help narrow the possible targets but would still be an overapproximated set at compile time (which would limit attack surface reduction). OCA opts to solve this at runtime. The instrumentation passes the function pointer to the runtime library, which then maps the appropriate page(s).

Using function pointers’ runtime values may require OCA to break its own rule of disallowing instrumentation inside of loops since one of its targets could be an encompassed function. When an indirect call is inside of a loop, OCA must still instrument it since its target may not be known outside the loop. On the other hand, as mentioned, repeatedly executing instrumentation and runtime library code inside of a loop can drastically degrade performance. Such cases are handled via optimization (discussed in Section 4.3).

Basic pseudocode for the compiler instrumentation pass is shown in Algorithm 1. It shows two functions, run_on_nonencompassed_func and run_on_func, which are hooks called by the pass manager on non-encompassed functions and all functions, respectively. The pseudocode shows the general logic for how decks are selected and inserted. Each deck needs only one key piece of runtime information, namely a unique ID that is generated statically for each loop or function. At runtime the library maps this parameter to a set of functions and their corresponding pages in memory. (In the case of indirect calls, the only difference is that the runtime target address is used instead of a statically known ID.) Several details not shown in the pseudocode that should be noted include: the insertion of the deck teardown calls; the insertion of an initialization call at program start; construction of functions’ static reachability; and construction of the encompassed function set.

OCA’s compiler instrumentation supports non-trivial C and C++ behavior. Though the details are unimportant, it is important to stress that the approach is general. Some of these features include the following: In addition to handling LLVM IR’s call instructions, it must also handle invoke instructions and therefore landing pads. It handles external libraries that take and then invoke a callback to the application. It handles recursion. It handles C++ destructors that can be invoked when an exception is thrown (via __cxa_throw). It handles libc_nonshared.a. It handles start-up C++ code before main. It handles signal handlers, including atexit and on_exit.

### 4.1.2 Linking

At the end of the compiler pass, OCA outputs a custom linker script. Intuitively, the goal of the linker script is to separate functions into different pages so that marking 1 function as RX does not “activate” unrelated functions (i.e. make them and their gadgets available for use). Two examples are helpful for understanding this stage, and we refer again to the xz callgraph example in Figure 3.

Figure 3: Simplified callgraph from the xz data compression application. This illustrates 4 types of edges, each of which requires different handling by the instrumentation pass.

Figure 3 depicts a sub-callgraph from SPEC CPU 2017’s xz, a data compression application [11]. It illustrates all but the indirect case. Each node in the figure is a function, and each edge is a call. Only the call from print_info_adv to msg_filters_to_str (dashed, green) is inside of a loop. The set of encompassed functions is therefore \{msg_filters_to_str,uint32_to_optstr\}, and the set of non-encompassed functions is \{main,msg_filters_show,print_info_adv\}. Instrumentation will be treated as follows:

1. The solid-black edges from main require single decks.
2. The dashed-green edge from print_info_adv requires a loop deck; instrumentation will be at the loop pre-header that dominates the call to msg_filters_to_str.
3. The solid-blue edge from msg_filters_show requires a reachable deck.
4. The dotted-red edge from msg_filters_to_str will have no instrumentation, because it is encompassed.
Algorithm 1 Pseudocode for OCA’s compiler pass.

```plaintext
function RUN_ON_NONENCOMPASSES_FUNC(func)
    for instr in func do
        if instr.IS_LOOP_BODY() then
            continue
        end if
        if instr.IS_LOOP_START() then
            INSERT_DECK(LOOP, instr.id)
        end if
        if instr.IS_DIRECT_CALL() then
            target ← instr.GET_CALL_TARGET()
            if target.IS_ENCOMPASSES() then
                INSERT_DECK(REACHABLE, target.id)
            else
                INSERT_DECK(SINGLE, target.id)
            end if
        end if
    end for
end function

function RUN_ON_FUNC(func)
    for instr in func do
        if instr.IS_INDIRECT_CALL() then
            target_addr ← instr.GET_FUNC_PTR()
            INSERT_DECK(INDIRECT, target_addr)
        end if
    end for
end function
```

Thus, the full callgraph’s deck sets $D_j \ast$ are as follows:

- $D_k.S_1 = \{\text{print\_info\_adv}\}$
- $D_k.S_2 = \{\text{msg\_filters\_show}\}$
- $D_k.L = \{\text{parse\_block\_header}, \text{msg\_filters\_to\_str}, \text{uint32\_to\_optstr}\}$
- $D_k.R = \{\text{msg\_filters\_to\_str}, \text{uint32\_to\_optstr}\}$

We follow a similar logic as in Example 1. Any of these functions can arbitrarily belong to the same system page at runtime. Thus, without any enforcement via the linker script, `print_info_adv` and `msg_filters_show` could again reside in the same page. Similarly, `parse_block_header` can occupy the same system page as either of the functions in $D_k.R$. That is, invoking `msg_filters_to_str` from `msg_filters_show` at runtime could inadvertently activate the gadgets in `parse_block_header`. The solution is to again rely on the fact that OCA instrumentation ensures that each deck will be mapped RX independently at runtime, and to leverage the custom linker script to avoid this security penalty. The custom linker script should separate the intersection $L \cap R = \{\text{msg\_filters\_to\_str}, \text{uint32\_to\_optstr}\}$ into its own disjoint set. Thus, the full disjoint sets $D_j \ast$ are as follows:

- $D_j.S_1 = \{\text{print\_info\_adv}\}$
- $D_j.S_2 = \{\text{msg\_filters\_show}\}$
- $D_j.L.1 = \{\text{parse\_block\_header}\}$
- $D_j.I.LR = \{\text{msg\_filters\_to\_str}, \text{uint32\_to\_optstr}\}$

—where $D_j.I.LR$ is the disjoint set formed by $L \cap R$ and $D_j.L.1$ is the disjoint set formed by $L \setminus (L \cap R)$. Each of these disjoint sets will be page-aligned by the custom linker script.

The pseudocode for creating these disjoint sets is shown in Algorithm 2. The algorithm begins with the “deck sets.” A deck set corresponds directly to 1 of the 4 types of decks: the function of a single deck forms a singleton; the functions of a loop deck form their own set; any encompassed function that can be called from some non-loop path has itself and any reachable functions as part of a set; and any functions that have their addresses taken and can be invoked by some indirect call form a set with their statically reachable callees. The algorithm begins with a list of these sets and then iterates, attempting to separate functions into their own disjoint sets, if possible. To find the disjoint sets, each pair of the decks is intersected with each other. If the intersection is non-null, those shared members are removed from the pair of decks and form their own disjoint set. This pairwise intersection-removal process is repeated until no more disjoint sets can be formed. Once the disjoint sets are known, each one is assigned its own page-aligned section in the linker script.

Example 1: Recall that `print_info_adv` and `msg_filters_show` are single decks. Without any enforcement via the linker script, these two functions could arbitrarily belong to the same system page at runtime. If OCA allowed this, then mapping one of these functions RX could inadvertently map the other RX, which exposes more code surface and therefore hurts overall security. But because these are different decks, the OCA instrumentation guarantees that each will be mapped RX independently at runtime. Thus, the custom linker script can safely separate these single decks to different page-aligned sections. That is, the decks $S_1 = \{\text{print\_info\_adv}\}$ and $S_2 = \{\text{msg\_filters\_show}\}$ will each form their own disjoint set. Each disjoint set is page-aligned by the custom linker script.

Example 2: We now consider the rest of the callgraph in Figure 3 and expand it by one node (see the updated callgraph in Figure 6 in Appendix Section A.1). In $xz$, there is in fact a function named `parse_block_header` between `print_info_adv` and `msg_filters_to_str` in its callgraph. It is on the same path and still part of the interprocedural loop.
Algorithm 2 Pseudocode for creating disjoint sets for OCA’s custom linker script.

```plaintext
function CREATE_DISJOINT_SETS(deck_sets)
  disjoint_sets ← ∅
  while `deck_sets.EMPT() do
    A ← deck_sets[0]
    tmp ← deck_sets[1 :]
    deck_sets ← ∅
    for B in tmp do
      I ← A ∩ B
      A_A ← A \ I
      if |I| == 0 then
        deck_sets.PUSH(B)
        continue
      end if
      if |A_A| == |B_I| == 0 then
        continue
      end if
      if |B_I| > 0 then
        deck_sets.PUSH(B_I)
      end if
      deck_sets.PUSH(I)
      A ← A_A
    end for
    disjoint_sets.PUSH(A)
  end while
  return disjoint_sets
end function
```

4.2 Runtime component

The runtime support is exposed as a library to the application. It is responsible for enabling and disabling pages by marking them RX or RO. The API is shown in Listing 2. These align directly with the 4 types of decks mentioned previously (single, loop, reachable, and indirect), plus library initialization. Although not shown in the listing, these API calls also have a corresponding deck teardown call to remap the relevant pages RO (and initialization has a corresponding destroy call).

Two important steps are necessary for library initialization. The first is identifying the binary’s base address. Function offsets are known at build-time, but at runtime OCA still needs to determine the text section’s base address. The second step is to protect all of the text pages by marking them RO. This happens at the start of main, and main is left RX.

When the program is running, any of the 4 main API endpoints may be invoked. `deck_single` takes as argument the function ID of an incoming callee. The runtime uses the ID to look up the actual page address of this function, and it marks it as RX. When that function returns, `deck_single_end` will mark the page as RO. `deck_reachable` is similar. It takes a callee function ID as its only argument. Because the callee is an encompassed function, though, all statically reachable functions must be marked RX, as well. Because these are compile-time known, the runtime library only needs to issue a map lookup to find which set of functions to mark RX for that particular callee.

`deck_loop` takes a single loop ID parameter as its argument. A unique loop ID is assigned by the compiler to each interprocedural, outermost loop in the program. When this runtime endpoint is invoked, it is a simple lookup to find which functions are part of that loop, and to mark them RX.

`deck_indirect` takes a runtime address as its argument. This is mapped by the library to the corresponding function, in order to determine whether that function is an encompassed or non-encompassed function. If it is encompassed, then the library leverages its own `deck_reachable` support for that function. Similarly, non-encompassed functions are handled by the `deck_single` support.

OCA maintains a reference counter for the text pages. When a function is needed, the reference count for each of its pages is incremented; when that function is no longer needed, the reference count for each of its pages is decremented. Whenever a page’s count changes from 0 to 1, the page must be marked RX; and whenever a page’s count changes from 1 to 0, it can be marked RO again. We define the set of pages at runtime with reference counts greater than 0 as the available pages. Adding a deck at runtime will either increase the cardinality of the available pages (if new pages are needed), or have no effect on its size (if all needed pages are already available); the opposite holds for removing a deck.

Listing 2: OCA’s runtime API.
```
int deck_init(void);
int deck_single(int callee_func_id);
int deck_loop(int loop_id);
int deck_reachable(int callee_func_id);
int deck_indirect(long long callee_addr);
```

4.3 Optimizations

We discuss one performance and one security optimization that we implemented in OCA. We call these two optimizations indirect deck caching (IDC) and stack cleaning (SC), respectively. Refer to Section 5.4 for their evaluation.

The most important performance optimization in OCA is reducing the overhead of indirect decks inside of loops. Instrumenting inside loops and just before indirect calls can reduce jump targets to a single function, but this comes at the expense of runtime overhead. To improve this, a combination of inlining and caching can eliminate nearly all overhead.

When any loop execution encounters an indirect call at runtime, the IDC optimization enforces an inlined check against the function pointer address. If the function pointer address is already cached by the runtime library, then the jump target’s page(s) must already be RX, and the application can proceed without any library call overhead. If not,
then the application pays some performance cost for invoking `deck indirect` on that iteration: The secure runtime library makes the page(s) available and caches the function pointer address in a hashmap. Note that the inlined code in the application is only for reading from the map, and the cache is cleared on loop exit. The intuition is that the combination of inlining the check, where the hardware’s branch predictor can be effective, and using a hashmap, where library call overhead is eliminated and the system’s memory cache can be leveraged, should severely limit extra cycles inside of loops.

For security, the SC optimization attempts to further reduce the attack surface. The general idea is to destroy decks in the call stack, and to only reconstruct them when returning up the call chain. This raises a question as to how much of the call stack should be “clean,” i.e. marked RO and unavailable. For example, SC could ensure only the last 4 decks are available on the call stack. Currently OCA only implements SC for single decks, and it imposes the strictest depth of 2. Thus, when `deck single` is invoked, SC only allows the current function and its upcoming callee to be mapped. Similarly, when `deck single end` is invoked, SC only allows the current function and its parent to be mapped.

5 Evaluation

We perform experiments on the SPEC CPU 2017 suite [11], 10 programs from the GNU coreutils package [10], and the nginx web server v1.20.1 [43]. We perform all experiments on a commodity desktop running an AMD Ryzen 7 1800X with 32GB RAM on Ubuntu 18.04 LTS. Our compiler is based on LLVM 11.0.0. Unless stated otherwise, OCA’s results are with the IDC and SC optimizations enabled (see Section 4.3). Here we present our performance and security findings. Our evaluation focuses on the following questions:

1. How much slowdown does an application incur when using the OCA framework? What is the code growth due to linker optimization for segregating function sets?
2. What is the gadget reduction for applications using OCA?
3. How effective are the optimizations in terms of performance and security?
4. In addition to general gadget reduction, can OCA break real gadget chains in the benchmarks and real-world applications to be able to stop gadget-based attacks?

5.1 SPEC CPU 2017

SPEC CPU 2017 [11] is a staple suite for CPU-bound performance benchmarking, making it useful for stressing the performance of binaries running under OCA. It also includes a diverse group of applications that give us insight into OCA’s effect on security, too. We use C and C++ applications from the suite, which are used in a variety of domains, including route planning, discrete event simulation, video compression, alpha-beta tree search, molecular dynamics, and ray tracing.

The performance results are reported in Figure 4. We compile and run a baseline version of each benchmark, optimized at -O3. Then we recompile and run the benchmark with OCA. The worst-case slowdown is 11% for omnetpp. Both imagick and perlbench have slight speedups, which can happen in instrumentation-based works that affect memory alignment [30, 45, 48, 68]; OCA also modifies function layout across pages which may play a role. The average slowdown across SPEC is 4%.

In comparison, BlankIt achieves 18% overhead on SPEC CPU 2006 by debloating libraries (not the application). Razor achieves 1.7% overhead on average on SPEC CPU 2006, with a worst-case of 16%. Piece-wise adds only negligible load-time overheads but deals only with libraries and not whole applications. Thus, we find OCA’s 4% average slowdown (on whole applications) in SPEC CPU 2017 to be reasonable compared with existing approaches.

Calculating gadget reduction for OCA is more complicated than in pre-runtime techniques like Razor, where the binary is trimmed, and the gadget reduction is measured and recorded offline. In OCA, however, the gadgets that are available to an attacker change dynamically during runtime based on which pages OCA has mapped RX, i.e. based on the available pages (discussed in Section 4.2).

To calculate these values, we start with the total number of gadgets in the application. This is the sum of ROP, JOP, COP, and special-purpose gadgets (reported by ROPGadget [52]):

\[ T = R + J + C + S \]  

(1)

The reduction of total gadgets for some set of available pages AP during runtime is given by the equation:

\[ \text{reduction}_{\text{AP}} = \frac{T - T_{\text{AP}}}{T} \times 100 \]  

(2)
Table 2: SPEC CPU 2017 total gadget reduction as a percentage (higher is better).

| Application | Min | Max | Avg |
|-------------|-----|-----|-----|
| perlbench   | 51.1| 98.8| 68.4|
| mcf         | 25.6| 68.4| 54.0|
| namd        | 75.4| 94.0| 88.5|
| parest      | 76.1| 99.8| 94.6|
| povray      | 36.6| 97.4| 53.0|
| lbm         | 47.9| 62.9| 57.4|
| omnetpp     | 52.4| 98.4| 79.1|
| xalancbmk   | 58.9| 99.6| 72.8|
| x264        | 17.2| 99.9| 32.5|
| blender     | 73.9| 99.8| 98.5|
| deepsjeng   | 24.4| 68.2| 64.9|
| imagick     | 39.3| 99.4| 88.7|
| leela       | 54.6| 87.7| 84.2|
| nab         | 68.6| 91.9| 86.6|
| xz          | 51.7| 94.9| 74.1|
| AVERAGE     | 50.2| 90.7| 73.2|

—where \( T \) is the total number of gadgets in the baseline application, and \( T_{AP} \) is the total number of gadgets in that available page set. Note, to calculate \( T_{AP} \), this is equivalent to Equation (1) for that specific available page set. For example, \( R_{AP} \) is calculated as:

\[
R_{AP} = \sum_{ap} R_{ap}
\]

—which is the sum of ROP gadgets \( R \) over each available page \( ap \) in the available page set \( AP \). Thus, the average reduction over all available page sets is equivalent to:

\[
\text{avg} \_ \text{reduction} = \frac{\sum \text{reduction}_{AP}}{\text{num} \_ \text{APs}}
\]

To capture these reduction metrics, we first enable OCA’s logging and run the program under all test inputs. This dumps the available pages on every OCA API call to a log file. Next, we scan every log line and identify the gadgets across each available page set. Each set of available pages is considered “equal” to another for the purposes of gadget-counting. The set of available pages that has the least proportion of reduced gadgets is marked as the “minimum reduction,” and vice versa for maximum reduction. The average reduction is the average proportion of gadgets reduced by each set of available pages over the logs.

The total gadget reductions are reported in Table 2. Each benchmark has three columns for minimum, maximum, and average reduction in total gadgets. OCA achieves an average of 73.2% total gadget reduction across all SPEC CPU 2017 benchmarks. The totals gadget reductions are representative of the individual (ROP, JOP, COP, and special-purpose) results, which is expected. For example, OCA’s average reduction of ROP gadgets specifically is 77.3%. OCA is not designed for reducing or favoring any particular type of gadget. Because it works at function and page granularity, it has a very similar reduction across all gadget types.

Direct comparisons are difficult because of differences in technique or reporting. Piece-wise reduces total gadgets by an average of 72.88% on SPEC CPU 2006 benchmarks for musl-libc. BlankIt reports an average of 97.8% ROP gadget reduction on SPEC CPU 2006 benchmarks for all libraries (and using glibc). Razor reports 68.19% code reduction (not gadgets) for applications in SPEC CPU 2006. Thus, OCA’s SPEC security result appears to be similar to the other application-focused technique. Razor, without sacrificing soundness. Compared with the library-only techniques, OCA appears to reduce applications’ gadgets equally as well as the load-time technique (Piece-wise), but not as thoroughly as an on-demand runtime technique (BlankIt).

5.2 GNU coreutils

We measure our technique on a subset of GNU coreutils. This package contains roughly 100 tools, including grep, mkdir, and rm. These utilities are relevant to software debloating for several reasons, including their real-world ubiquity, and that they have a history of CVEs. Chisel and Razor also report results for coreutils that OCA can compare against.

The authors of Razor made their tool available [13], so we use the same application versions and inputs as them. Their inputs were designed to cover the same functionality tested by Chisel. We use only the test inputs, as we do not require any training. The number of inputs per benchmark ranges between 17-40, and the number of options that any given input may exercise ranges from 1-7 (see [49] for more details).

Runtime overheads are negligible for coreutils. Every test completes in under 1 second and is trivially performant. In contrast, SPEC CPU 2017 tests each take 3-10 minutes.

Table 3: GNU coreutils total gadget reduction as a percentage (higher is better).

| Application | Min | Max | Avg |
|-------------|-----|-----|-----|
| bzip2       | 42.7| 78.8| 70.8|
| chown       | 88.3| 97.3| 95.9|
| date        | 95.0| 97.5| 96.9|
| grep        | 65.0| 90.9| 82.8|
| gzip        | 34.7| 75.7| 64.6|
| mkdir       | 90.4| 96.6| 94.5|
| rm          | 88.4| 98.7| 96.9|
| sort        | 79.0| 91.9| 90.5|
| tar         | 49.0| 86.8| 83.4|
| uniq        | 93.0| 96.0| 95.4|
| AVERAGE     | 72.5| 91.0| 87.2|

As with SPEC, we present the total gadget reduction numbers (see Table 3). The average decrease across coreutils
is 87.2%. The worst-case scenario occurs at one point during gzip, where only 34.7% of the application’s gadgets are unavailable. The best case occurs for rm at 98.7%. In comparison, Razor and Chisel achieve 61.9% and 85.1% ROP gadget reduction on coreutils, respectively. Thus, OCA compares well against two other state-of-the-art techniques that reduce ROP gadgets in application code.

### 5.3 nginx

nginx is by some metrics the most popular web server today [62]. Besides serving web content, it is also deployed frequently as a reverse proxy or load balancer. As a common multitool in today’s web infrastructure, security is a real concern for nginx. It is multithreaded and multiprocessed, unlike the SPEC benchmarks and coreutils applications we have evaluated, and this can break or stress frameworks. nginx’s performance is a critical factor in certain deployments, so it is important that any security techniques not interfere too heavily with it. It is also recently evaluated by another debloating technique, BlankIt, which will serve as a good comparison point. For all these reasons, we choose to evaluate nginx with OCA.

We faithfully reproduce the security and performance experiments described in the BlankIt evaluation [48]. We use the same nginx workload generator, wrk, which runs 12 threads in parallel and creates 400 concurrent connections to the server. For the performance experiment, there are 4 inputs: the home page of Wikipedia, and 3 randomized binaries of 1MB, 10MB, and 100MB. The performance experiment includes 2 separate tests. In the first test, wrk requests the Wikipedia home page for 3s, then 30s, then 300s. The experiment tests that nginx can serve a normal-sized page (80KB) under high load for extended periods of time without degrading. In the second test, wrk requests the 1MB binary for 30s, then the 10MB binary for 30s, and finally the 100MB binary for 30s. This experiment tests that as the request size scales, there is still no degradation. For the security experiment, only the home page of Wikipedia is used. wrk makes requests for 30s.

The performance result is shown in Figure 5. The slowdown is reported as the transfer/sec degradation, normalized against the baseline. OCA achieves 1.023x slowdown on average. Table 4 shows the total gadget reduction on nginx. The average is 80.3% (an improvement over SPEC, but less than the reduction for coreutils).

| Application | Min | Max | Avg |
|-------------|-----|-----|-----|
| nginx       | 50.3| 95.3| 80.3|

In comparison, BlankIt averages 1.047x runtime overhead and 98.9% ROP gadget reduction on nginx’s libraries. As

Figure 5: Transfer/sec degradation for nginx using OCA (normalized against a baseline).

with SPEC, OCA outperforms BlankIt at runtime but with less ROP gadget reduction. BlankIt copies needed library code into place before use and zeroes it out after use. This accounts for BlankIt’s higher gadget reduction, and also explains why, despite only being used on libraries, BlankIt is slower than OCA’s page-mapping scheme.

### 5.4 Benefits from Optimization

We evaluate two optimizations, indirect deck caching (IDC) and stack cleaning (SC) (see Section 4.3). We also evaluate the effect of the modified linker script used by OCA.

IDC can be broken into two parts: inlining a check inside loops to see if a function pointer is already mapped; and waiting until loop exit to disable the function pointer’s page(s). We found that both of these steps are critical to good performance. For example, SPEC’s xz application suffers from over 10x slowdown without IDC. Disabling the page(s) after loop exit reduces the slowdown to under 8x. Inlining the check for the mapped pages and using caching reduces the slowdown to only 1.05x. The inlined check is the major factor across benchmarks. For example, perlbench, x264, and povray have over 5x, 2.5x, and 1.3x slowdown, respectively, without it.

In terms of security, IDC and the custom linker script (CLS) are highly beneficial. OCA achieves 28.3% total gadget reduction on SPEC without either of these techniques. Applying only CLS, this improves to 43.1%. Applying only IDC, it improves to 66.4%. Together they achieve the 73.2% total reduction reported previously. IDC’s benefit to attack surface reduction is a strong argument against fully static treatment of function pointers, which leaves too many gadgets available at loop headers. The SC optimization, however, seems to have little effect. On SPEC, we measured only negligible improvement in total gadget reduction (<1%). Because it is conservative (being applied only on single decks) and does not negatively impact performance, we have left it active in
cases where security could benefit. The result suggests that SC would need to be expanded in future work to handle other types of decks.

5.5 Binary Size Growth

We measured the binary size increase over every application. There is 2.9x increase across SPEC, 1.5x across coreutils, and 1.8x for nginx. In absolute terms, the modified binaries are 17MB, 1.2MB, and 7.1MB on average. Thus, for these applications, the binaries are still reasonably sized, despite the growth, and the performance measurements have confirmed that this has not adversely affected runtime. The improvement of coreutils over SPEC is due to fewer disjoint sets in coreutils. All binaries in coreutils have fewer than 200 sets, whereas the majority in SPEC have 200 or more.

The worst-case growth without any custom linking would assign a page-aligned section to every function in the program (i.e., every disjoint set would be a singleton). We estimate this case, lower-bounding it by ignoring weak function symbols in the baseline applications. Our custom linking script improves over lower-bounded, worst-case growth by 1.8x, 2.7x, and 1.3x for SPEC, coreutils, and nginx, respectively.

5.6 Breaking ROP Gadget Chains

Gadget reduction is a common metric in security-focused debloating works [24, 48, 49, 51, 57], but it is still difficult to draw certain conclusions. For example, even with 90% total gadget reduction, a significantly large enough binary could still have an enormous number of gadgets in absolute terms. Alternatively, 50% reduction may be good, provided that it removes a critical class of gadgets needed by any useful chain. Thus, gadget reduction is a useful yardstick, but we also need to investigate whether an attack surface reduction technique actually removes attacks.

Ropper [54] is an open source tool that can automatically build gadget chains given some constraints. It allows us to automatically test if we can successfully build a ROP gadget chain that spawns /bin/sh via an execve syscall. We use Ropper to identify which binaries from all our previous experiments have this ROP gadget chain. Then we test every binary over all test inputs with OCA and check every available page set for the ROP gadget chain.

Table 5 shows the prevalence of the ROP chain across all applications’ text sections. (blender’s analysis times out, so it is not included.) Recall from Listing 1 that this ROP chain requires a gadget to store values in memory, a sequence of gadgets to set up the syscall arguments, and a gadget to invoke the execve syscall. These correspond to Table 5’s W-W-W, Args, and Syscall columns, respectively. (W-W-W stands for a write-what-where gadget.) If all of these components of the gadget chain are present for an application, then the end-to-end exploit exists (a checkmark in the “E2E exploit” column).

Out of 25 baseline applications, 8 of them have the full ROP gadget chain. 12 have an incomplete chain, and 5 have no components of the chain.

We analyze every available page set over all inputs across all applications with Ropper and find that OCA does not allow the ROP chain under any set of dynamic decks. The analysis includes 5,390 unique dynamic deck sets (with 12,718 dynamic execution count) over all applications and inputs.

We analyze these chains and their broken counterparts under OCA to validate the result and understand how OCA is achieving 100% effectiveness at breaking this ROP chain. We show an analysis of sort to explain how the syscall opcode manifests and why OCA removes it in all test runs. The syscall opcode in x86_64 is 0x0F05. Ropper identifies a syscall opcode in sort within a jump instruction. An objdump with 1 line of context is shown here:

```
4276ae:  41 b0 79  mov $0x79,%r8b
4276b1:  e9 0f 05 00 00 jmpq 427bc5
4276b6:  80 fa 8a  cmp $0x7b,%d1
```

The jump target is strftime_case_.254+0xfb5. The e9 byte is the jump opcode, and 0f 05 00 00 is the addend to a base (which in little endian is 0x00000050F). The base is calculated from the program counter (of the next instruction), which is 0x4276b6 in this case. Thus the jump target is 0x4276B6 + 0x050F, which yields 0x427BC5. Notice that it is by chance that the jump target offset happens to be 0x050F. This code is part of the strftime_case function. This is a custom version of strftime from glibc that is built into coreutils. As such, it would not have been handled by load-time techniques like Piece-wise or library-only techniques like BlankIt. In these experiments, this function’s page is never mapped RX for sort, and thus OCA eliminates this syscall opcode.

6 Related Work

Researchers have developed multiple defenses to deal with code reuse attacks. One of the most prominent techniques is control flow integrity (CFI) [1, 18]. CFI limits forward control flow transfers to legal edges in the control flow graph (CFG) and callgraphs. It is typically paired with a shadow stack, which restricts backward control flow transfers to the legal target on the stack. CFI has a rich history, addressing a variety of scenarios, contexts, and attacks [4, 12, 16, 22, 25, 28, 34, 38, 44, 60, 61, 67].

CFI defends against “control data attacks,” but it does not directly protect against “non-control data attacks” [9], which corrupt data but not code pointers. Like control data attacks, they can also be used to launch code reuse attacks. Castro et al. introduced data-flow integrity (DFI) in 2006 [7], which helps defend against both control and non-control data attacks. DFI
Table 5: Execve-to-shell ROP chain availability across all baseline applications.

| App       | W-W-W | Args | Syscall | E2E exploit |
|-----------|-------|------|---------|-------------|
| perlbench | ✓     | ✓    | ✓       | ✓           |
| mcf       | ✓     | ✓    | ✓       | ✓           |
| namd      | ✓     | ✓    | ✓       | ✓           |
| parest    | ✓     | ✓    | ✓       | ✓           |
| povray    | ✓     | ✓    | ✓       | ✓           |
| lbm       | ✓     | ✓    | ✓       | ✓           |
| omnetpp   | ✓     | ✓    | ✓       | ✓           |
| xalancbmk | ✓     | ✓    | ✓       | ✓           |
| x264      | ✓     | ✓    | ✓       | ✓           |
| deepsjeng | ✓     | ✓    | ✓       | ✓           |
| imagick   | ✓     | ✓    | ✓       | ✓           |
| leela     | ✓     | ✓    | ✓       | ✓           |
| nab       | ✓     | ✓    |         |             |
| xz        | ✓     | ✓    |         |             |
| bzip2     |       | ✓    |         |             |
| chown     | ✓     | ✓    |         |             |
| date      |       | ✓    |         |             |
| grep      | ✓     | ✓    |         |             |
| gzip      | ✓     | ✓    |         |             |
| mkdir     | ✓     | ✓    |         |             |
| rm         | ✓     | ✓    |         |             |
| sort      | ✓     | ✓    |         |             |
| tar       | ✓     | ✓    |         |             |
| uniq      | ✓     | ✓    | ✓       | ✓           |
| nginx     | ✓     | ✓    | ✓       | ✓           |

tracks values and ensures they are not updated improperly between moves (i.e. that they adhere to the original data flow of the program).

Memory protection techniques attempt to protect memory itself, i.e. prevent the corruption as opposed to just the exploit. Softbound+CETS provides complete memory safety at 116% runtime overhead without hardware support [15, 40, 41]. Code pointer integrity [30] provides protection on only code pointers, which keeps overhead reasonable, but leaves non-control data attacks out of scope.

Despite these advancements, current state-of-the-art still has its shortcomings. There are numerous examples of how to bypass CFI (e.g. [19]) or what its limits are (e.g. [6]). In fact, recent work [33] thoroughly categorizes the shortcomings of several CFI systems, and which they broadly characterize as: imprecise analysis methods, improper runtime assumptions, unprotected corner code, unexpected optimization, incorrect implementation, mismatched specification, and unintended targets. For example, πCFI [45], which leverages Modular CFI (MCFI) [44], uses structural equivalence for type checking. This would treat void * the same as char *, for instance. This imprecision allows extra indirect targets for the attacker to use. µCFI [25] cannot protect against code pointer reuse and VTable attacks. OS-CFI [27] fails to protect against indirect calls that are optimized for indirect calls.

As mentioned, non-control data attacks are out of scope for CFI and CPI. Though DFI provides some protection for such attacks, data flow analysis is still an overapproximation and therefore insecure. Furthermore, the seminal DFI paper achieved 44-103% runtime overhead on SPEC CPU 2000 [7], which is generally considered too slow for practical use. Similarly, Softbound+CETS, despite offering full memory protection, suffers from these high overheads.

Thus, traditional techniques like CFI for guarding against control flow hijacking have had their shortcomings acknowledged. DFI, CPI, and memory protection also have certain disadvantages that make them in some way incomplete (e.g. because their defense is incomplete, their scope is limited, or their overheads are not acceptable for a certain task). In short, opportunities for attackers to launch code reuse attacks exist today and are expected to continue. This has led to recent work that attempts to reduce software’s attack surface.

Software deboating and attack surface reduction are an orthogonal solution to approaches such as CFI or DFI, and can be seen as a security hardening technique. In addition to [24, 48, 49, 51], other prominent work includes feature-based techniques. Slimium [50] debloats Chromium features based on a static-, dynamic-, and heuristic-based analysis. Koo et al. take a configuration-driven approach to remove feature-specific code [29], achieving 77% debo late on nginx. Trimmer [57] is another technique that takes as input user configuration and uses it to drive the deboating process.

Lastly, software engineering researchers have worked on deboating, but the focus has not traditionally been on security. For instance, [35–37, 63–66] leverage deboating to improve performance, and [3, 14, 21, 39] use it to reduce code size.

7 Conclusion

We present OCA, an attack surface reduction technique that works on full programs and can enable may-use code on-demand. It achieves state-of-the-art gadget reduction results without compromising soundness or requiring training or user inputs. Total gadget reduction across SPEC CPU 2017, GNU coreutils, and the nginx server average 73.2%, 87.2% and 80.3%, respectively. In our performance experiments, the runtime slowdown on SPEC is 4% and negligible for GNU core-util s; the transfer/sec degradation for nginx is only 2%. In an additional study over these applications, we show that for all test inputs, OCA eliminates the presence of a ROP chain that spawns a shell via execve. Based on these results and the generality of the approach, we find OCA to be a promising technique for attack surface reduction.
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A Appendix

This appendix includes: An expanded callgraph of xz, which is helpful for the example in Section 4.1.2, and a brief discus-
sion of future work.

A.1 Expanded \texttt{xz} Callgraph

![Expanded callgraph from the \texttt{xz} data compression application. This expanded graph includes one additional node, \texttt{parse_block_header}.]

Figure 6: Expanded callgraph from the \texttt{xz} data compression application. This expanded graph includes one additional node, \texttt{parse_block_header}.

A.2 Future Work

There are several interesting directions that we did not explore with OCA. The first is leveraging Intel’s MPX [26] hardware for secure hashing for the IDC optimization. In [27], the authors show how to repurpose MPX to securely hash and store metadata. Because support for it is being removed from GCC, it is available for tools like OCA to trial. This could improve IDC’s performance further, and in fact, it opens the possibility of applying this caching optimization to other types of decks inside of loops. If a loop exercises only a sparse subset of its interprocedural, statically reachable function set, then this could potentially have large security benefits.

We also identified the Intel Memory Protection Keys (MPK) [26] hardware primitive as a way to drastically reduce the runtime’s overhead for remapping pages. OCA’s runtime could potentially replace \texttt{mprotect} calls with a single WRPKRU instruction. The libmpk library [46] makes this secure and simpler, and reports 8.1x speedup for equivalent \texttt{mprotect} calls. A related insight is that the granularity for building and tearing down decks does not necessarily need to be at the function and page level. For that matter, the granularity does not need to be fixed. By leveraging MPK, for example, a deck could be composed of a set of basic blocks within a loop, and good performance might still be achievable.