Making Code Re-randomization Practical with MARDU

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Abstract—Defense techniques such as Data Execution Prevention (DEP) and Address Space Layout Randomization (ASLR) were the early role models preventing primitive code injection and return-oriented programming (ROP) attacks. Notably, these techniques did so in an elegant and utilitarian manner, keeping performance and scalability in the forefront, making them one of the few widely-adopted defense techniques. As code re-use has evolved in complexity from JIT-ROP, to BROP and data-only attacks, defense techniques seem to have tunneled on defending at all costs, losing-their-way in pragmatic defense design. Some fail to provide comprehensive coverage, being too narrow in scope, while others provide unrealistic overheads leaving users willing to take their chances to maintain performance expectations.

We present MARDU, an on-demand system-wide re-randomization technique that improves re-randomization and refocuses efforts to simultaneously embrace key characteristics of defense techniques: security, performance, and scalability. Our code sharing with diversification is achieved by implementing reactive and scalable, rather than continuous or one-time diversification while the use of hardware supported eXecute-only Memory (XoM) and shadow stack prevent memory disclosure; entwinning and enabling code sharing further minimizes needed tracking, patching costs, and memory overhead. MARDU’s evaluation shows performance and scalability to have low average overhead in both compute-intensive (5.5% on SPEC) and real-world applications (4.4% on NGINX). With this design, MARDU demonstrates that strong and scalable security guarantees are possible to achieve at a practical cost to encourage deployment.

I. INTRODUCTION

Present day computing continues to trudge through a challenging jungle of memory corruption vulnerabilities with no clear endgame in sight. Early in the journey through the attack landscape jungle, code injection attacks [16], [18], [53] were subsided with the introduction of simplistic yet effective techniques like Data Execution Prevention (DEP) [52], [63].

This quickly refocused efforts against stronger adversaries such as code re-use. Code re-use attacks, like return-oriented programming (ROP) [74] and ret-into-libc [75], utilize a victim’s program code against itself. Innocent code snippets, called gadgets, are repurposed to construct gadget-chains, the equivalent of a malicious payload and initiated via memory corruption vulnerabilities. To combat code weaponization and locating these gadgets in the first place, coarse-grained ASLR [23], [24], [54], [79], was created to obfuscate the base address of an executable code section while keeping all code intact. Because of ASLR’s cost-to-benefit ratio and upholding of code sharing, this pushed for position-independent code to become the default compilation method for applications to utilize it.

Finding that coarse-grained ASLR could be bypassed by a single memory disclosure vulnerability, ASLR moved to fine-grained randomization via randomizing the executable code region on the scale of instructions [47], [57], basic blocks [34], [67], or memory pages [22]. However, instruction displacement [57], as well as in-place code randomization (IPR) [67], do not have full coverage of all gadgets present, leaving some gadgets completely exposed. Oxymoron [22] focused on protecting code pointers replacing all code references with unique labels and accessing functions via indexing into a protected table (RaTTle). To note, Oxymoron is one of the few defenses that is scalable, supporting code sharing and not using background processes for tracking or patching code pointers.

Even if code was randomized at very fine granularity, indirect memory disclosures (via stack and heap) could still be used to reveal code pages during runtime as in just-in-time return-oriented programming (JIT-ROP), leaving fine-grained ASLR ineffective. By collecting addresses and repeatedly exploiting memory disclosure vulnerabilities, JIT-ROP springboards the attacker to reveal the entire valid executable code region. Additionally, Snow et al. [77] presented a JIT-ROP capable of reading JIT’ed code within an application boundary and construct ROP gadget chains on the fly. Isomeron [34] was the only fine-grained ASLR technique capable of preventing both traditional ROP as well as JIT-ROP, by randomly switching execution between two versions of program code making it unpredictable which version will be chosen.

eXecute-only Memory (XoM) was then implemented to more thoroughly guard memory and restrict access to finding gadgets. XoM was either “resilient” (e.g., Heisenbyte [78]) via destructive code reads or “resistant”, completely preventing reading of the code region as in Readactor [32] by using Extended Page Tables (EPT). While effective against JIT-ROP, XoM at that time was in its infancy. Tagging and virtualizing memory via the use of EPT made any memory access expensive and hog system resources, leaving fewer resources for the user. It would not be until recently that XoM would have hardware support via Intel Memory Protection Keys (MPK) [50] and similar protection in ARMv7-M [20]. Though code could not be directly read via XoM, it could still be deduced via inference attacks and de-randomization techniques. BROP [26] utilized generalized stack crash reading to de-randomize ASLR and leak enough gadgets to launch a code reuse attack. Other code inference attacks involve crash-less reads [38], [44], allocation oracles [66], or zombie gadgets via shared code reloading [76]. These attacks showed that randomizing once is simply not sufficient even if protected under XoM; code remains static thereafter and still vulnerable to any memory leak.

At this point, attacks became more fierce, intimidating defense techniques to defend by any means necessary and get out of the jungle alive. Re-randomization enables code
to become too volatile to reliably craft attacks, rendering any leaked knowledge or weaponized code stale and incorrect, thus thwarting attacks. Currently, it is assumed that most attacks are carried out remotely or require I/O system calls to engage. To counter this, some works opt to trigger via a time interval, ideally shorter than the network round-trip latency [29], [37], [88], while others use known “code re-use relevant” events such as \texttt{fork()} [61] or I/O system calls [25], [84]. However, this assumption does not paint the entire picture of the ROP attack landscape. Defenses that utilize a \textit{threshold} whether via time, such as Shuffler [88] and CodeArmor [29], or via leaked-data write() to only full-function code reuse (e.g., Shuffler [88] and CodeArmor [29], or via leaked-bytes as in ReRanz [84], are vulnerable to inevitably faster low-profile attacks as attacks continue to evolve. Therefore it can be reasoned that using an interval is simply not a comprehensive metric and that the concept of using a \textit{threshold} is a liability.

TASR [25] does not rely on a threshold, expecting attackers usage of I/O system calls like \texttt{write()}. While they do prevent remote JIT-ROP that use I/O system calls, TASR cannot prevent memory disclosure within the application boundary [72]. These re-randomization techniques forgo expensive execute-only memory, and instead pair randomization with protecting code pointers. Early re-randomization works such as RuntimeASLR [61] and TASR [25] use mutable code pointer approaches to ideally perturb a significant amount of live code, but this comes with an equally significant performance cost (e.g., 30-50% in TASR [84], [88]) associated with pointer tracking and patching at re-randomization, severely limiting the effective re-randomization frequency possible.

Immutable code pointer approaches such as CodeArmor, Shuffler, and ReRanz are much more lightweight in terms of tracking and patching. Although CodeArmor uses segmentation (e.g., \%gs) with offsetting, this technique still allows for ROP gadgets as well as function pointers to be reached (e.g., $f+o = \%gs:f+o$, where $f$ is an immutable code pointer and $o$ is an offset). Trampoline and indirection tables, used by Shuffler and ReRanz, respectively, constrain this loophole to only full-function code reuse (e.g., allow only the case of $f+o = f'$, where $f'$ is another function in the trampoline); completely eliminating full-function code reuse and data-oriented programming [49] is ongoing research.

Regrettably, no current re-randomization techniques take into consideration the scalability of their approaches; support for code sharing has been forgotten, and the prevalence of multi-core has excused the reliance on per-process background threads. In short, re-randomization is not as impenetrable and competitive as initially thought. Current work has shown that making a secure, practical, and scalable ROP defense technique is challenging. Even if recent defenses have made some headway through the jungle, most still lack effective comprehensiveness in security for the system resource demands they require in return (both CPU and memory); these factors are prime showstoppers for deployment.

In this paper, we introduce MARDU to refocus defense technique design, showing that it is possible to embrace the core fundamentals of performance and scalability, while ensuring comprehensive security guarantees. MARDU builds on insight that thresholds like time intervals as in Shuffler and CodeArmor or leaked-data-amount in ReRanz are a security loophole and a performance shackle in re-randomization; MARDU does not rely on a threshold whatsoever in its design. This lets MARDU completely side-step the no-win trade-off between security and performance. Instead, MARDU borrows the event trigger design but pairs it with XoM violations. Using XoM, MARDU provides complete prevention of JIT-ROP protecting against both variations of remote JIT-ROP as well as local JIT-ROP, compared to TASR which can only defend against the former, with almost zero overhead by using Intel MPK XoM. MARDU also combines XoM with trampolines by covering them from read access while also completely decoupling the function entry and the function body in memory; unlike in CodeArmor, this makes it impossible to infer and obtain ROP gadgets in the middle of a function from a leaked code pointer.

MARDU keeps performance and scalability at its forefront. MARDU does not require expensive code pointer tracking and patching like TASR, nor does MARDU incur significant overhead from continuous re-randomization triggered by over-conservative time intervals or benign I/O system calls as in Shuffler and ReRanz, respectively. Additionally, while TASR shows a very practical average overhead of 2.1%, it has been reported by Shuffler [88] and ReRanz [84] that TASR’s overhead against a more realistic baseline (not using compiler flag -D) is closer to 30-50% overhead. Finally, MARDU is designed to both support code sharing and not require the use of any additional system resources (e.g., background threads as used in numerous works [25], [29], [37], [84], [88]) with the help of Linux kernel memory management and leveraging its own calling convention. To summarize, this paper makes the following contributions:

- \textbf{ROP attack & defense analysis.} Our background §II describes the four prevalent ROP attacks that challenge current works, including JIT-ROP, code-inference, low-profile, and code pointer offsetting attacks. In addition, we describe the bottom-line security implications for each attack. With this, we classify and exhibit current state-of-the-art defenses standings on three fronts: security, performance, and scalability. Our findings show most defenses are not as secure or as practical as expected against current ROP attack variants.

- \textbf{MARDU defense framework.} We present the design of MARDU in §IV, a comprehensive ROP defense technique capable of addressing all currently known ROP attacks.

- \textbf{Scalability and shared code support.} MARDU creates and uses a new calling convention in order to be able to both leverage a shadow stack and minimize the overhead of pointer tracking. This calling convention is also what enables shared code (e.g., libraries) to be even more secure, able to be re-randomized by any host process and maintain the integrity for the rest of the entire system. To the best of our knowledge, MARDU is the first framework capable of this.

- \textbf{Evaluation & open source prototype.} We implement a prototype of MARDU based on LLVM and Linux kernel. }
evaluate and analyze MARDU in §VI with compute-intensive benchmarks as well as real-world applications. MARDU’s overhead for compute-intensive benchmarks is 5.5% on average (geometric mean) and its worst-case overhead is 18.3%. We will open source MARDU for the community to explore other defense mechanisms and build upon our work.

II. CODE LAYOUT (RE-)RANDOMIZATION

In this section, we present a background on existing code layout re-randomization techniques. To help understand the attack and defense arms race, we classify code randomization techniques into two categories: load-time randomization and continuous re-randomization.

To start with, we describe two attacks, A1 (JIT-ROP) and A2 (BROP), aimed to defeat load-time code randomization. Next, we describe how continuous re-randomization techniques defy A1 and A2 by analyzing design elements of the techniques. Specifically, we categorize continuous re-randomization techniques by its re-randomization triggering condition (i.e., either based on the system call history or timing threshold) and the semantics of storing code pointers (i.e., tracking pointers or use an indirect, immutable function index). The reason for focusing on those two categories is that these design elements greatly affect the security, performance, and scalability of techniques. Finally, we compare and contrast each technique regarding three aspects, security, performance, and scalability. Particularly for illustrating attack resilience, we present two attacks, A3 (low-profile JIT-ROP) and A4 (code pointer offsetting), to which existing re-randomization techniques are susceptible. Regarding scalability, we report the requirement of additional CPU usage for re-randomization and whether shared code layout is possible among multiple processes.

In summary, Table I illustrates the characteristics of each defense technique by randomization category, attack resilience, and performance and scalability factors, and we describe these in detail in the following.

A. Attacks against Load-time Randomization

Load-time code randomization techniques suffer from attacks A1 (JIT-ROP) and A2 (BROP, etc.). In the following, we describe characteristics of techniques and attacks.

1) Load-time Randomization without XoM: Code layout randomization techniques so-called coarse-grained ASLR [79] or fine-grained ASLR [22], [31], [34], [47], [48], [54], [57], [67], [86], depending on the granularity of layout randomization, fall into this category of code layout randomization. These techniques randomize the code layout only once, usually when code is loaded into memory. After code is loaded and shuffled, its layout never changes during the lifetime of the program. The following attack can defeat the security guarantee of load-time randomization techniques.

A1: Just-in-time ROP (JIT-ROP). An attacker with arbitrary memory read capability may launch JIT-ROP [77] by interactively performing memory reads to disclose one code pointer. This disclosure can be used to then leap frog and further disclose other addresses to ultimately learn the code contents in memory. Any load-time code randomization technique that does not protect code from read access including fine-grained ASLR techniques is susceptible to this attack.

| Security implications: | Techniques failing to protect code from read access allows code-reuse attacks to be launched regardless of code randomization granularity. |
|------------------------|--------------------------------------------------------------------------------------------------------------------------------|

2) Load-time Randomization with XoM: In response to A1 (JIT-ROP), several research projects protect code from read access via destructive read memory [78] or execute-only memory [21], [27], [29], [32], [33], [41], [70], [78], [87]. Systems with destructive read [78] allow code pointer leaks to occur, but trigger intended localized code corruption once a read attempt is made, such that code-reuse attacks following the read will fail to execute the expected code by the attacker. Systems with execute-only memory (XoM) [21], [27], [29], [32], [33], [41], [70], [78], [87] aim to fundamentally block all read attempts of program code by removing read permissions from the code area. Applying these techniques prevent attackers from gaining knowledge about code contents, and thereby, nullifying A1. However, leaving the code layout fixed after load-time randomization makes these techniques susceptible to the following attack.

A2: Blind ROP (BROP) and code inference attacks. Even with protecting code from read access (i.e., XoM), load-time randomizations still are susceptible to BROP [26] and/or other code content inference attacks [69], [76]. Although these attacks do not read code directly, attackers may accumulate information about the code contents by conducting probing on the code many times because code layout will never change after it is loaded and shuffled. In particular, BROP is a clone-probing attack that infers code contents via observing differences in execution behaviors such as timing or program output. Other attacks [69], [76] defeat destructive code read defenses [78], [87] by inferring code contents from a small fraction of a code read and then weaponizing inferred code.

| Security implications: | Maintaining a fixed layout over crash-probing or read access to code allows inferring code contents indirectly, and thereby, attackers can still learn the code layout and launch code-reuse attacks. |

B. Continuous Re-randomization Defeating A1 & A2

In response to A1 (JIT-ROP) and A2 (BROP, etc.), continuous re-randomization techniques [25], [29], [37], [42], [61], [84], [88] aim to defeat attacks by continuously shuffling code (and data) layouts at runtime to make information (code or code addresses) leaks or code probing done before shuffling useless.

To illustrate the internals of re-randomization techniques in a nutshell, we describe the core design elements of re-randomization by categorizing them into two, based on their design elements: 1) Re-randomization triggering condition and 2) Code pointer semantics.

Re-randomization triggering condition. Existing continuous re-randomization techniques trigger their randomization based on the following two conditions.
### Table I: Classifications of ASLR-based code-reuse defenses. Gray highlighting emphasizes the attack (A1-A4) that largely invalidated each type of defense. ○ indicates the attack is blocked by the defense. (attack-resistant). × indicates the defense is vulnerable to that attack. ▲ indicates the attack is not blocked but is still mitigated by the defense (attack-resilient). ✓ indicates the defense meets performance/scalability requirements. × indicates the defense is unable to meet performance/scalability requirements. N/A in the column A3 indicates that the attack is not applicable to the defense due to lacking of re-randomization, and N/T in the column Performance indicates that either SPEC CPU2006 or perlbench is not tested. Specifically in the column Perf., ▲ indicates that the defense cannot prevent the JIT-ROP attack within the application boundary that does not use system calls; in the column A4, ▲ indicates that an attack may reuse bothROP gadgets and entire functions while ▲ indicates that an attack can only reuse entire functions. † Note that in TASR the baseline performance is a binary compiled with -0g, necessary to correctly track code pointers. Previous work [84], [88] reported performance overhead of TASR using regular optimization (-02) binary is ≈30-50%. MARDU provides strong security guarantees with low performance overhead and good system-wide scalability compared to existing re-randomization approaches.

| Types       | Defenses                  | Security         | Performance       | Scalability       |
|-------------|---------------------------|------------------|-------------------|-------------------|
|             |                           | Gran. | A1 | A2 | A3 | A4 | Perf. | Avg. | Worst | Code Sharing | No Addl. | Process |
| Load-time ASLR | Fine-ASLR [31], [47], [48], [54], [57], [67], [86] | Fine          | ☑  | ☑  | N/A | ☑  | ✓     | 0.4% | 6.4% | ☑  | ☑  | ✓     |
|             | Oxypron [22]               | Coarse         | ☑  | ☑  | N/A | ☑  | ✓     | 2.7% | 11%  | ☑  | ✓     |
|             | Isomeron [34]              | Fine           | ☑  | N/A | ☑  | ▲   | ✓     | 19%  | 42%  | ☑  | ☑  | ✓     |
| Load-time+NoM | Readactor/Readactor++ [32], [33] | Fine          | ☑  | N/A | ☑  | ▲   | ✓     | 8.4% | 25%  | ☑  | ☑  | ✓     |
|             | LR 2 [27]                  | Fine           | ☑  | N/A | ▲   | ✓   | ✓     | 6.6% | 18%  | ☑  | ☑  | ✓     |
|             | kR’X [70]                  | Fine           | ☑  | N/A | ▲   | ✓   | ✓     | 2.32%| 12.1%| ☑  | ☑  | ✓     |
| Re-randomization | TASR [25]             | Coarse         | ☑  | ☑  | N/A | ☑  | ×     | 2.1% | 10.1%| †  | ×     | ×     |
|             | ReRanz [84]                | Fine           | ☑  | ☑  | ▲   | ✓   | ▲     | 5.3% | 14.4%| ×   | ×     | ×     |
|             | Shuffler [88]              | Fine           | ☑  | ▲  | ▲   | ×   | ✓     | 14.9%| 40%  | ☑  | ×     | ×     |
|             | CodeArmor [29]             | Coarse         | ☑  | ☑  | N/A | ☑  | ×     | 3.2% | 55%  | ☑  | ×     | ×     |
| Our Approach | MARDU                      | Fine           | ☑  | ☑  | ☑  | ▲   | ✓     | 5.5% | 18.3%| ✓   | ✓     | ✓     |

### Timing:
- Techniques [29], [88] shuffle the layout periodically by setting a timing window for layout randomization. For example, Shuffler [88] triggers re-randomization every 50 msec, and CodeArmor [29] can set re-randomization period as low as 55 μsec.

### System-call history:
- Techniques [25], [61], [84] shuffle the layout based on the history of the program’s previous system call invocations, e.g., after invoking fork() [61] or when write() (leak) is followed by read() (exploit) [25], [84].

### Code pointer semantics.
Existing continuous re-randomization techniques use the following three different types of code pointer semantics.

- **Code address as code pointer:** Code pointers store the actual addresses. In this case, leaking a code pointer lets the attacker have knowledge about code address. Therefore, techniques in this category [25], [42] require tracking of code pointers (or all pointers) at runtime, which is computation expensive, to update their values after randomizing the code and data layout. For instance, TASR [25] shows very high performance overhead (30-50%) especially for I/O-intensive applications, such as web servers [84], [88].

- **Function trampoline address as code pointer:** Code pointers store a function table index [88] or the address of a function trampoline [84]. This design avoids the expensive pointer tracking in order to enhance the performance of re-randomization techniques. Instead of tracking and updating code pointers, techniques in this category setup a function table, which stores all function addresses of the program, and store an index of the table in the code pointer to refer a function. After re-randomization and re-locating the code layout, the techniques update only the function table while all code pointers remain immutable. With this design, leaking a code pointer will tell the attacker only the semantics of referring to a function (i.e., function index in the trampoline) but not about the code layout.

- **An offset to the code address as code pointer:** Code pointers store an offset from the (randomized) base address of the layout. This design is also intended for avoiding pointer tracking by having an immutable offset from the random version address for referring to a function, as in CodeArmor [29]. At re-randomization, updating code layout only requires updating the random version base address, and does not require any update of pointers. With this design, leaking a code pointer will tell the attacker the offset to select a function no matter how the code layout is randomized.

### C. Attacks against Continuous Re-randomization
Continuous re-randomization techniques suffer from two attacks (A3 and A4) that we define in this section.

#### A3: Low-profile JIT-ROP
This is class of attacks does not trigger re-randomization either by completing the attack quickly or without requiring I/O system calls. As the trigger for layout re-randomization, Existing defenses utilize one of timing [29], [37], [42], [88], amount of transmitted data by output system calls [84], or I/O system call boundary [25] as the trigger for layout re-randomization. Therefore attacks within the application boundary, such as code-reuse attacks in Javascript engine where both information-leak followed by control-flow hijacking attack may conclude faster than the re-randomization timing threshold or not interact with I/O system call, can bypass these triggering conditions. The code layout may remain unchanged within the given interval, and thereby, an attacker may launch JIT-ROP to unchanged code layout.
A4: Code pointer offsetting. Even with re-randomization, techniques might be susceptible to a code pointer offsetting attack if code pointers are not protected from having arithmetic operations applied by attackers \[25, 29\]. An attacker may trigger a vulnerability to apply arithmetic operations to an existing code pointer. For example, suppose a code pointer \( p \) points to a function \( f() \), and altering the value of \( p \) by adding an offset \( o \) could make \( p \) point another code address \( p+o \). Particularly, in techniques directly using code address \[25\] or code offset \[29\], \( p+o \) could be even a ROP gadget in \( f() \) if the attacker knows the gadgets offset \( o \) beforehand. Ward et al. \[85\] has recently demonstrated that this attack is possible against TASR.

**Security implications:** Maintaining a fixed code layout across re-randomizations and not protecting code pointers lets attackers perform arithmetic operations over pointers, allowing launching other ROP gadgets.

### III. Threat Model and Assumptions

We build MARDU based on the following assumptions.

**Attacker’s capability:**
- **Arbitrary read/write.** Attackers can perform arbitrary memory read/write (if the address is readable/writeable) in a target process by exploiting software vulnerabilities in the victim program. With this capability, attackers may launch attacks A1–A4 on an unprotected system.
- **Local brute-force attacks.** We assume that all attack attempts are run in a local machine. In this regard, attacks may be performed any number of times within a short time period (e.g., within a millisecond). This assumption gives the capability of launching attacks A1–A4 without triggering re-randomization in prior systems \[25, 29, 42, 84, 88\].

**System and trusted computing base:**
- **XoM (R\(\oplus\)X) and DEP (W\(\oplus\)X).** We assume that the userspace of the system does not have any memory region that is both readable and executable. Likewise, we assume that the userspace of the system does not have any memory region that is both writable and executable.
- **Trusted hardware and no physical access.** We assume all hardware is trusted and attackers do not have physical access. Particularly, we trust Intel Memory Protection Keys (MPK) \[50\], a mechanism that provides eXecute-Only Memory (XoM), and we regard attacks to CPU (side-channel attacks, e.g., Spectre \[55\], Meltdown \[59\]) to be out of scope.
- **Trusted kernel and program loading.** We trust the OS kernel and the loading/linking process of the program (\texttt{execve()}, \texttt{1d-\text{}linux}, etc.), thus attackers cannot intervene to perform any attack before the program starts.

### IV. MARDU Design

We begin by the design overview of MARDU \((\S IV-A)\) and then detail MARDU compiler \((\S IV-B)\) and kernel \((\S IV-C)\).

**A. Overview**

This section presents the overview of MARDU, along with its design goals, challenges, and outlines its architecture.

1) **Goals:** Understanding how the attack landscape and existing mitigations fit together, our goal in designing MARDU is to shore up the current state-of-the-art to enable a practical code randomization. More specifically, our design goals are as follows:

**Security.** No prior solutions provide a comprehensive defense against existing attacks (see \(\S II\)). Systems with only load-time ASLR are susceptible to leaking code-content (A1) and letting attackers infer code-content (A2). Systems applying re-randomization are still susceptible to low-profile attacks (A3) and code pointer offsetting attacks (A4). MARDU aims at either defeating or significantly limiting the capability of attackers in launching code-reuse attacks spanning from A1 to A4 to provide best-effort security against existing attacks.

**Performance.** Many prior approaches \[29, 32, 33, 88\] demonstrate decent runtime performance in average cases \(< 10 \%\), e.g., \(< 3.2 \%\) in CodeArmor); however, they also show evidence of a few scenarios that are remarkably slow (i.e., > 55 %, see listed numbers in Worst column in Table 1). We design MARDU to run with an acceptable average overhead \((\approx 5 \%)\) with minimal performance outliers across a variety of application types.

**Scalability.** Most proposed exploit mitigation mechanisms have overlooked the impact of required additional system resources, such as memory or CPU usage, which we consider a scalability factor. This is crucial for applying a defense system-wide, and is even more critical when deploying the defense in the pay-as-you-go pricing Cloud. Oxymoron \[22\] is the only defense that allows code sharing of randomized code. No advanced re-randomization defenses support code sharing thus they require significantly more memory. Additionally, most re-randomization defenses \[29, 84, 88\] require per-process background threads, which not only cause additional CPU usage but also contention with the application process. As a result, approaches requiring per-process/thread background threads show significant performance overhead as the number of processes increases. For example, Shuffler \[88\] shows around 55% performance overhead when four NGINX workers (plus four Shuffler threads) run on two cores with 50 ms shuffling interval.

Therefore, to apply MARDU system-wide, we design MARDU to not require significant additional system resources, for instance, additional processes/threads or significant additional memory.
2) Challenges: It is challenging to achieve all of aforementioned goals. One naive approach is to mix all good defenses from existing approaches but such an approach fails to meet the goal because requirements for enabling each defense could conflict. Hence, we list challenges in achieving our goals in the following.

Tradeoffs in security, performance, and scalability. An example of the tradeoff between security and performance is having fine-grain ASLR with re-randomization. Although such an approach can defeat A4, systems cannot apply such protection because the re-randomization must finish quickly to meet the performance goal and also defeat A3. An example of the tradeoff between scalability and performance is having a dedicated process/thread for performing re-randomization. However, this results in a drawback in scalability by requiring more CPU time in the entire system. Therefore, a good design must find a breakthrough to meet all of aforementioned goals.

Conflict in code-diversification vs. code-sharing. Layout re-randomization requires diversification of code layout per process, and this affects the availability of code-sharing. The status quo is that code sharing cannot be applied to any existing re-randomization approaches, and this makes the defense unable to scale to protect many-process applications. Although Oxymoron [22] enables both diversification and sharing of code, it does not consider re-randomization, nor use a sufficient randomization granularity (page-level), which is insufficient against A4.

3) Architecture: We design MARDU to make a breakthrough beyond tradeoffs in security, performance, and scalability, which satisfies all three aspects and becomes practical. We introduce our approach for each aspect below:

Scalability: Sharing randomized code. MARDU manages the cache of randomized code in the kernel, which is capable of being mapped to multiple userspace processes and is not readable from userspace. Thus, it does not require any additional memory.

Scalability: System-wide re-randomization. Since code is shared between processes in MARDU, per process randomization, which is CPU intensive, is not required; rather a single process randomization is sufficient for the entire system. For example, once a worker process of NGNIX server crashes, it re-randomizes all its mapped executables (e.g., libc.so) upon exit. This re-randomizes all processes using the same executables (e.g., libc.so of all processes, including other workers of the NGNIX server, will be immediately re-randomized).

Fig. 1: Overview of MARDU

Performance: Immutable code pointers. The above described design decisions for scalability also help reduce performance overhead. In addition, MARDU neither tracks nor encrypts code pointers so it does not mutate code pointers upon re-randomization. While this design choice minimizes performance overhead, other security features (e.g., XoM, trampoline, and shadow stack) in MARDU ensure the comprehensive ROP defense.

Security: Detecting suspicious activities. MARDU considers any process crash or code probing attempt as a suspicious activity. MARDU’s use of XoM makes any code probing attempt trigger process crash and system-wide re-randomization. Therefore, MARDU counters direct memory disclosure attacks as well as code inference attacks requiring initial code probing [69], [76]. To implement XoM, we use Intel MPK [50] so our XoM design does not impose any runtime overhead unlike virtualization-based designs.

Security: Preventing code & code pointer leakage. In addition to system-wide re-randomization, MARDU is designed to minimize the leakage of code and code pointers. Besides XoM, we use three techniques. First, MARDU applications always go through a trampoline region to enter into or return from a function. Thus, only trampoline addresses are stored in memory (e.g., stack and heap) while non-trampoline code pointers remain hidden. MARDU does not randomize where the trampoline region is so MARDU does not need to track and patch code pointers in memory upon re-randomization. Second, MARDU performs fine-grained function-level randomization within an executable (e.g., libc.so) to completely disconnect any correlation between trampoline addresses and code addresses. This provides high entropy (i.e., roughly n! where n is the number of functions), so it is not feasible to succeed BROP [26] without any crash. Also, unlike re-randomization approaches that rely on shifting code base addresses [25], [29], [61], MARDU is not susceptible to code pointer offsetting attack (A4). Finally, MARDU stores return addresses—precisely, trampoline addresses for return—in a shadow stack; the shadow stack stores only return addresses and is hidden under a segmentation register in x86. This design makes stack pivoting practically infeasible.

Design overview. As shown in Figure 1, MARDU is composed
of compiler and kernel components. The MARDU compiler replaces call and ret instructions with functionally equivalent jmp’s to the trampoline region and generates code to store return addresses in a shadow stack. MARDU compiler generates PC-relative code so randomized code can be shared by multiple processes. Also, the compiler generates and attaches additional metadata to binaries for efficient patching of PC relative addressing code upon (re-)randomization. The compiler separates data from code pages to prevent false-positive XoM violations, from MARDU applications attempting to read inter-mixed data in protected code regions.

The MARDU kernel is responsible for choreographing the runtime when a MARDU executable is launched. The kernel extracts and loads the executable’s compiler-generated metadata into a cache to be shared by multiple processes. This data is then used by MARDU for first load-time randomization as well as re-randomization. The randomized code is cached and shared by multiple processes; while allowing sharing, each process will get a different random virtual address space for the shared code. MARDU kernel prevents read operations of the code region including the trampoline region using XoM so trampoline addresses do not leak information about non-trampoline code. Whenever a process crashes (e.g., XoM violation), MARDU kernel re-randomizes all associated shared code so all relevant processes are re-randomized to thwart an attacker’s knowledge immediately.

B. MARDU Compiler

MARDU compiler generates a binary able to 1) hide its code pointers, 2) share its randomized code among processes, and 3) run under XoM. To this end, MARDU uses its own calling convention using a trampoline region and shadow stack.

1) Code Pointer Hiding: Trampoline. MARDU hides code pointers without paying for costly runtime code pointer tracking. The key idea for enabling this is to split a binary into two regions in process memory: trampoline and code regions (as shown in Figure 2 and Figure 3). A trampoline is an intermediary call site that moves control flow securely to/from a function body, protecting the XoM hidden code region. There are two kinds of trampolines: call and return trampolines. As their names imply, a call trampoline is responsible for forwarding control flow from an instrumented call to the code region function entry, while a return trampoline is responsible for returning control flow semantically to the caller. Each function has one call trampoline to its function entry, and each call site has one return trampoline returning to the following instruction of the caller. Since trampolines are stationary, MARDU does not need to track code pointers upon re-randomization because only stationary call trampoline addresses are exposed to memory.

Shadow stack. Unlike the x86 calling convention using call/ret to store return addresses on the stack, MARDU instead stores all return addresses in a shadow stack and leaves data destined for the regular stack untouched. Effectively, this protects all backward-edges. An instrumented call pushes a return trampoline address to the shadow stack and then jumps to a call trampoline; an instrumented ret directly jumps to the return trampoline address at the current top of the shadow stack. The base address of the MARDU shadow stack is randomized by ASLR and is hidden in segment register %gs, which cannot be modified in userspace and will never be stored in memory. Therefore, it is infeasible to know the shadow stack base address without causing a program crash. We additionally reserve one register, %rbp, to use exclusively as a stack top index of a shadow stack in order to avoid costly memory access.

Running example. Figure 2 is an example of executing a MARDU-compiled function foo(), which calls a function bar() and then returns. Every function call and return goes through trampoline code which stores the return address to a shadow stack, of which base address is hidden in register %gs. The body of foo() is entered via its call trampoline 1. Before foo() calls bar(), the return trampoline address is stored to the shadow stack–each call site has one return trampoline returning to the next instruction of the call site. Control flow then jumps to bar()’s trampoline 2, which will jump to the function body of bar() 3. bar() returns to the address in the top of the shadow stack, which is the return trampoline address 4. Finally, the return trampoline returns to the instruction following the call in foo() 5.

2) Enabling Code Sharing among Processes: PC-relative addressing. To enable sharing, MARDU compiler generates PC-relative (i.e., position-independent) code so code can be shared amongst processes that load the same code in different virtual addresses. The key challenge here is how to incorporate PC-relative addressing with randomization. MARDU randomly places code (at function granularity) while trampoline regions are stationary. This means any code using PC-relative addressing must be correspondingly fixed up once its randomized location is decided. In Figure 2, all jump targets between

![Fig. 2: Illustrative example executing a MARDU-compiled function foo(), which calls a function bar() and then returns.](image-url)
the trampoline and code, denoted in yellow rectangles, are PC-relative and must be fixed. Also, all data addressing instructions are PC-relative (e.g., accessing global data, GOT, etc.) and also must be fixed.

**Fixup information for patching.** With this policy, it is necessary to keep track of these instructions to patch them properly during runtime. To make the runtime patching process simple and efficient, MARDU compiler generates additional metadata into the binary that describes exact locations for patching and their file-relative offset. This fixup information makes the patching as simple as just adjusting PC-relative offsets for given locations within analyzing instructions (see Figure 3). Since displacement in PC-relative addressing is 32 bits in size in x86-64 architecture, ±2 GB is the maximum offset from the %rip supported by this addressing mode. We elaborate on the patching process in §IV-C2.

**Supporting a shared library.** A call to a shared library is treated the same as a normal function call to preserve the code pointer hiding property; that is, MARDU refers to the call trampoline for the shared library call via PLT/GOT. It first calls the PLT (Procedure Linkage Table) via a trampoline, which jumps to an external function whose address is not known at link time and left to be resolved by the dynamic linker. The result of dynamic symbol resolution is a function address in the call trampoline of the external library, and it is stored in a GOT (Global Offset Table) for caching. While MARDU does not specifically protect GOT, we assume that GOT is already protected using MPK [36], [62].

**Overhead of sharing.** MARDU’s sharing mechanism does not have noticeable runtime overhead as PC-relative code is already mandatory to enable ASLR. In addition, the overhead of runtime patching is negligible because MARDU avoids “stopping the world” when patching the code to maintain internal consistency compared to other approaches.

3) **Enabling Execute-Only-Memory:** Finally, to run code with XoM, MARDU compiler ensures code and data are segregated in different pages. Compilers sometimes intermingle data within .text code section as an optimization. However, if this data is attempted to be read during runtime, an XoM violation will be raised. As previous work [32] reported, we found that Clang intermingles code and data only for jump tables so we disable generating jump tables in .text section.

C. **MARDU Kernel**

MARDU kernel randomizes code at load-time and runtime. It maps already-randomized code, if it exists, to the address space of a newly fork-ed process. When an application crashes, MARDU re-randomizes all mapped binaries associated with the crashing process and reclaims the previous randomized code from the cache after all processes are moved to a newly re-randomized code. MARDU prevents direct reading of randomized code from userspace using XoM. MARDU initializes a shadow stack whenever clone-ing a task.

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1In this paper, a term task denotes both process and thread as the convention in Linux kernel.

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2We note that, for the unused region, we map all those virtual addresses to a single abort page that generate a crash when accessed to not to waste real physical memory and also detect potential attack attempts.

3We choose 2 GB because in x86-64 architecture PC-relative addressing can refer ±2 GB range from %rip.
a high entropy, MARDU kernel uses fine-grained randomization within the allocated virtual address region. Once the trampoline is positioned, MARDU kernel randomly places the non-trampoline code within the virtual address region (sized 2 GB); MARDU decides the random offset between the code and the trampoline regions. Once the code region is decided, MARDU permutes functions within the code region to further increase entropy. As a result, trampoline addresses do not leak information on non-trampoline code and an adversary cannot infer any actual codes’ location from the system information, /proc/<pid>/maps, as they will get the same mapping information for the entire 2 GB region. **Patching the randomized code.** After permuting functions, MARDU kernel patches instructions accessing code or data according to randomization. MARDU kernel patches `%rip` relative offsets in instructions. This patching process is trivial at runtime; MARDU compiler generates fixup location information in the binary and MARDU kernel re-calculates and patches PC-relative offsets of instructions according to the randomized function location. Note that patching includes control flow transfer between the trampoline and non-trampoline code and global data access (i.e., .data, .bss) as well as function calls to other shared libraries (i.e., PLT/GOT).

3) Randomized Code Cache: MARDU kernel manages a cache of randomized code. When a userspace process tries to map a file with executable permissions, MARDU kernel first looks up whether there exists a randomized code of the file in cache. If cache hits, MARDU kernel maps the randomized code region to the virtual address of the requested process. Upon cache miss, it performs load-time randomization as described earlier. MARDU kernel manages how many times the randomized code region is mapped to userspace. If the reference counter is zero and the memory pressure of the system is high, MARDU kernel evicts the randomized code. Thus, in normal cases without re-randomization, MARDU randomizes a binary file only once. In our implementation, the randomized code cache is associated with the `inode` cache. Thus, when the `inode` is evicted from the `inode` cache under severe memory pressure, its associated randomized code is also evicted.

4) Execute-Only Memory (XoM): We designed XoM based on Intel MPK [50]4. With MPK, each page is assigned to one of 16 domains (referred to as a protection key), which is encoded in a page table entry. Read and write permissions of each domain can be independently controlled through an MPK register. When randomized code is mapped to a userspace virtual address, we set the permissions of the corresponding page table entries to executable, which is in fact executable and readable, and assign code memory to the XoM domain. MARDU kernel configures the XoM domain to non-accessible (i.e., neither readable nor writable) so MARDU kernel can enforce execute-only permission with MPK. If an adversary tries to read XoM-protected code memory, MARDU kernel will raise SIGBUS and trigger re-randomization. Unlike EPT-based XoM designs [32], [78], our MPK-based design does not impose runtime overhead.

5) On-Demand Re-randomization: Triggering re-randomization. An unsuccessful probing of the attack causes the process to crash. Therefore, when a process crashes MARDU triggers re-randomization of all binaries mapped to the crashing process. Since MARDU re-randomization thwarts attacker’s knowledge (i.e., each attempt is an independent trial), an adversary must succeed in her first try without crashing, which is practically infeasible.

**Re-randomizing code.** Upon re-randomization, MARDU kernel first populates another copy of the code (e.g., libc.so) in the code cache and freshly randomizes it (Figure 4 1). MARDU places the trampoline code at the same location not to change trampoline addresses to avoid mutating code pointers but it randomly places the non-trampoline code (i.e., random offset in Figure 3 2) such that the new one does not overlap with the old one. Then, it permutes functions in the code. Thus, the re-randomized code is completely different from the previous one without changing trampoline addresses.

**Live thread migration without stopping the world.** The re-randomized code prepared in the previous step is not yet visible to userspace processes because it is not yet mapped to userspace virtual address space. To make it visible, MARDU first maps the new non-trampoline code to application’s virtual address space, Figure 4 2. Because the old trampoline code is still mapped, the new code is not reachable yet. Then, MARDU remaps the virtual address range of the trampoline code to the new trampoline code by updating corresponding page table entries 3. After this, the new trampoline code will transfer control flow to the new non-trampoline code so that any thread crossing the trampoline migrates to the new non-trampoline code without stopping the world.

**Safety reclaiming the old code.** MARDU can safely reclaim the code only after all threads migrates to the new code 4. MARDU uses reference counting for each randomized code to check if there is a thread accessing the old code. After the new trampoline code is mapped 5, MARDU sets a reference counter of the old code to the number of all runnable tasks 5 that map the old code. It is not necessary to wait for migration of non-runnable, sleeping task because it will correctly migrate to the newest randomized code region when it passes through the (virtually) static return trampoline, which refers to the new layout when it wakes up. The reference counter is decremented when a runnable task enters into MARDU kernel due to system call or preemption. When calling a system call, MARDU kernel will decrement reference counters of all code that needs to be reclaimed. When the task returns to userspace, it will return to the return trampoline and the return trampoline will transfer to the new code. When a task is preempted out, it may be in the middle of executing the old non-trampoline code. Thus, MARDU kernel not only decrements reference counters but also translates `%rip` of the task to the corresponding address in

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4As of this writing, Intel Xeon Scalable Processors [51] and Amazon EC2 C5 instance [17] support MPK. Other than the x86 architecture, ARM AArch64 architecture also supports execute-only memory [19].

5A task in a TASK_RUNNING status in Linux kernel.
the new code. Since MARDU permutes at function granularity, \%rip translation is merely adding an offset between the old and new function locations.

**Summary.** Our re-randomization scheme has three nice properties: time boundness of re-randomization, almost zero overhead of running process, and system-wide re-randomization. The re-randomization is guaranteed to finish at most within one scheduling quantum (e.g., 1 msec) once the newly randomized code is exposed. That is because MARDU migrates runnable tasks at system call and scheduling boundary. If another process crashes in the middle of re-randomization, MARDU will not trigger another re-randomization until the current randomization finishes. However, as soon as the new randomized code is populated, a new process will map the new code immediately. Therefore, the old code cannot be observed more than once. MARDU kernel populates a new randomized code in the context of a crashing process. All other runnable tasks only additionally perform reference counting or translation of \%rip to the new code. Thus, its runtime overhead for runnable tasks is negligible. To the best of our knowledge, MARDU is the first system to perform system-wide re-randomization allowing code sharing.

6) Shadow Stack: To hide the shadow stack location, MARDU first reserves a 2 GB of virtual memory space with abort pages and then chooses a base address to map the shadow stack. Additionally, no direct reference to the shadow stack is available in memory because MARDU accesses it via a dedicated register, \%gs, which will not disclose the base address of the shadow stack. As a result, an adversary must brute-forcingly guess its base address; any crash in such attempt will trigger re-randomization, which invalidates all prior information gained. When a new task is created (clone), the MARDU kernel allocates a new shadow stack and copies parent’s shadow stack to its child.

V. IMPLEMENTATION

We implemented MARDU on the Linux x86-64 platform. MARDU compiler is implemented using LLVM 6.0.0 and MARDU kernel is implemented based on Linux kernel 4.17.0 modifying 3549 and 4009 lines of code (LOC), respectively. We used musl libc 1.1.20 [13], which is a fast, lightweight C standard library implementation for Linux, because glibc cannot be compiled with Clang. We manually modified 164 LOC in musl libc to make assembly functions (e.g., math functions and atomic intrinsics) follow the MARDU calling convention by adding C wrapper functions so MARDU compiler can automatically identify and instrument them.

A. MARDU Compiler

**Trampoline.** MARDU compiler is implemented as backend target-ISA (x86) specific MachineFunctionPass. This pass instruments each function body as described in §IV-B.

**Re-randomizable code.** To force all instructions to use PC-relative addressing, MARDU compiler uses -fPIC. As an optimization of our trampoline design, MARDU compiler uses a register \%rbp for stack top index of a shadow stack. To force the compiler to relinquish the use of the register \%rbp, MARDU compiler uses -fomit-frame-pointer. To save space in PC-relative addressing of code and data, the x86 architecture provides various jmp instruction variants varying in offset size from 1, 2, or 4 bytes. In MARDU, to maximize entropy and be able to use the full span of memory within our declared 2 GB virtual address region, MARDU compiler uses -mrelax-all to force the compiler to always emit full 4-byte displacement in the executable. Since all PC-relative instructions have 4-byte wide displacement, MARDU kernel can freely place any function to within 2 GB address range without any restriction. To completely separate code and data for XoM, MARDU compiler disables jump tables in .text section using -fno-jump-tables.

B. MARDU Kernel

**Shadow stack.** Currently the maximum shadow stack size is set to 64 KB; when a task is created, MARDU kernel creates a 2 GB virtual address space region and randomly place its shadow stack in that region, guarded by inaccessible pages (to hide the shadow stack). To maximize performance, MARDU implements a compact shadow stack without comparisons [28].

**Secure random number generation.** To perform randomization, MARDU uses cryptographically secure random number generator in Linux kernel based on hardware random sources.
We choose SPEC CPU2006 over its newer version, SPEC which complies to MA RD U. We evaluate the performance overhead and web server. The SPEC benchmark suite has various realistic Assembly Code. M ARDU does not support inline assembly as it difficult to figure how to deal with module-level inline assembly as was present in musl libc; however, this could be resolved with further engineering. Instead, we manually added C wrapper functions so M ARDU compiler adds trampolines which complies to M ARDU calling convention. Setjmp and exception handling. M ARDU uses a shadow stack to store return addresses. Thus, functions such as setjmp, longjmp, and longjmp directly manipulate return addresses on stack are not supported by our current prototype. However, modifying these functions is straightforward because essentially our shadow stack is a variant of compact, register-based shadow stack [28]. C++ support. Our prototype does not support C++ applications since we do not have a stable standard C++ library that is musl-compatible. Therefore handling C++ exceptions and protecting vtables is out of scope.

VI. Evaluation

We evaluate M ARDU by answering these questions:

- How secure is M ARDU, when presented against current known attacks on randomization? (§VI-A)
- How much performance overhead does M ARDU impose, particularly for compute-intensive benchmarks? (§VI-B)
- How scalable is M ARDU in a real-world network facing server, particularly with concurrent processes, in terms of load time, re-randomization time, and memory used? (§VI-C)

Applications. We evaluate the performance overhead and scalability of M ARDU using SPEC CPU2006 and NGINX web server. The SPEC benchmark suite has various realistic compute-intensive applications (e.g., gcc) which are ideal to see the worst-case performance overhead of M ARDU. We tested all 12 C language benchmarks; we excluded C++ benchmarks as our current prototype does not support it. We choose SPEC CPU2006 over its newer version, SPEC CPU2017, because SPEC CPU2006 has been popularly used to show the performance overhead in many prior works. Input size ref was used for all benchmarks. To test performance and scalability of M ARDU on a complex, real-world application, we ran NGINX, which is a widely utilized web server. We configured NGINX as multi-process. We report the average of four runs.

Experimental setup. All programs are compiled with optimization -O2 and run on a 24-core (48-hardware threads) machine equipped with two Intel Xeon Silver 4116 CPUs (2.10 GHz) and 128 GB DRAM.

A. Security Evaluation

To evaluate the security of M ARDU, we first analyze the resiliency of M ARDU against existing attacker models against load-time randomization (A1–A2, §VI-A1) and continuous re-randomization (A3–A4, §VI-A2). Then, to illustrate the effectiveness of M ARDU for a wider class of code-reuse attacks beyond ROP, we present results of the residual attack surface analysis using the threat model of NEWTON [82] (§VI-A3).

M ARDU Security Summary:
- vs. A1: Execute-only memory blocks the attack.
- vs. A2: Re-randomization blocks any code inference via crash.
- vs. A3: Execute-only memory and a large search space (2 GB dummy mappings) block JIT-ROP and crash-resistant probing.
- vs. A4: Trampolines decouple function entry from function bodies blocking any type of code pointer offsetting; full function code reuse of exported functions remains possible.
- vs. NEWTON: the same as A4.

1) Attacks against Load-Time Randomization: Against JIT-ROP attacks (A1). M ARDU asserts permissions for all code areas (both code and trampoline regions) as execute-only (via XoM); thereby, an attacker with JIT-ROP capability cannot read code contents directly.

Against code inference attacks (A2). M ARDU blocks code inference attacks, including BROP [26], clone-probing [61], and destructive code read attacks [69], [76] via layout re-randomization triggered by an application crash or XoM violation. This mechanism effectively blocks A2 attacks by preventing attackers from accumulating indirect information because every re-randomization renders all previously gathered (if any) information regarding the code layout invalid.

Hiding shadow stack. Attackers with arbitrary read/write capability (A1/A2) may attempt to leak/alter shadow stack contents if its address is known. Although the location of the shadow stack is hidden behind the %gs register to prevent leakage of pointers, attackers may employ attacks that undermine this sparse-memory based information hiding [35], [44], [66]. To prevent such attacks, M ARDU reserves a 2 GB virtual memory space for the shadow stack (the same way M ARDU allocates code/library space) and then chooses a random offset to map the shadow stack; other pages in the remaining 2 GB space are mapped as an abort page that has no permissions. Even assuming if an attacker is able to identify the 2 GB region for the shadow stack using crash-less poking [44] or employing allocation oracles [66], they must also overcome the randomization entropy of the offset to get a valid address within this region; any incorrect probe will generate crash (due to abort pages), thereby thwarting the attack. Consequently, the probability of successfully guessing the location of any valid shadow stack address is roughly one in $2^{31}$, practically infeasible.

Entropy. M ARDU applies both function-level permutation and random start offset to provide high entropy to the new code layout. Specifically, M ARDU permutes all functions in each executable and applies a random start offset to the code area in 2 GB space for each randomization. Thus, randomization entropy depends on the number of functions in the executable and the size of a code region (i.e., $\log_2(n + 2^{31})$ where $n$ is the number of functions). To give an idea of how much
Entropy MARDU provides, we take an example of 470.1bm in SPEC CPU2006, a case which provides the minimum entropy in our evaluation. The program, which contains 26 functions and is less than 64 KB in size, has 119.38 bits entropy. Therefore, even for a small program, MARDU randomizes the code with significantly high entropy (119 bits) to render attacker’s success rate for guessing the layout negligible.

2) Attacks against Continuous Re-randomization: Against low-profile attacks (A3). MARDU does not rely on timing nor system call history for triggering re-randomization. As a result, neither low-latency attacks nor attacks without involving system calls are effective against MARDU. Instead, re-randomization is triggered and performed by any MARDU instrumented application process on the system that encounters a crash (e.g., XoM violation). Nonetheless, a potential A3 vector could be one that does not cause any crash during exploitation (e.g., attackers may employ crash-resistant probing [35], [39], [44], [56], [66]). In this regard, MARDU places all code in execute-only memory with 2 GB mapped region. Such a stealth attack could only identify multiples of 2 GB code regions and will fail to leak any fine-grained layout of code or code addresses stored in trampolines.

Against code pointer offsetting attacks (A4). For code pointers referring to call trampolines, attackers may attempt to launch an A4 attack by adding/subtracting an offset to the pointer. To defend against such an attack, MARDU decouples any correlation between trampoline function entry addresses and function body addresses (i.e., no fixed offset), so attackers cannot refer to the middle of a function for a ROP gadget without actually obtaining a valid function body address. Additionally, the trampoline region is also protected with XoM, thus attackers cannot probe it to obtain function body addresses to launch A4. MARDU limits available code-reuse targets to only exported functions in the trampoline region. We analyze the residual attack surface of MARDU in Table II.

3) Viable Attacks in MARDU: Attack analysis with NEWTON. To measure the boundary of viable attacks against MARDU, we present a security analysis of MARDU based on the threat model set by NEWTON [82]. In this regard, we analyze possible writable pointers that can change the control flow of a program (write constraints) as well as possible available gadgets in MARDU (target constraints), which will reveal what attackers can do under this threat model. In short, MARDU allows only the reuse of exported functions via call trampolines.

For write constraints, attackers cannot overwrite real code addresses such as return addresses and code addresses in the trampoline. MARDU only allows attackers to overwrite other types of pointer memory, e.g., object pointers and pointers to the call trampoline. For target constraints, attackers can reuse only the exported functions via call trampoline. Note that a function pointer is a reusable target in any re-randomization techniques using immutable code pointers [29], [84], [88]. Although MARDU allows attackers to reuse function pointers in accessible memory (e.g., a function pointer in a structure), such live addresses will never include real code addresses, such as a return address or real code address, and will be limited to addresses referencing call trampolines. Under these write and target constraints, inferring the location of ROP gadgets from code pointers (e.g., leaking code addresses or adding an offset) is not possible.

Residual attack surface. Table II presents potential code-reuse attack surface of programs in this evaluation and how much MARDU reduces such attack surface (i.e., possible code reuse targets in the program). For each program, we present the attack surface in three categories.

- Direct calls (# call) and indirect calls (# i-call), which may leak return addresses to the regular stack.
- ROP Gadgets (# gadgets), the main ingredient in constructing a code re-use payload by being chained together to make a valid attack. This data is obtained via running ROPgadget [73] on each benchmark binary.
- Function entries (# fn entry) that can be used for whole function re-use attack.

MARDU completely protects the first two categories (i.e., return address and ROP gadget) using shadow stack and XoM, respectively. Therefore, the remaining potential attack surface is function entry (i.e., call trampolines in MARDU). Evaluating SPEC, musl libc, and NGINX, MARDU reduces access to these sensitive fragments up to a max of 97.1% and 95.6% on average, constraining attackers to reuse only exported function entries.

B. Performance Evaluation

Runtime performance overhead with SPEC CPU2006. Figure 5 shows the relative performance overhead of SPEC with MARDU compared to the unprotected baseline, compiled with vanilla Clang. Overall, MARDU’s average overhead is comparable to the fastest re-randomization systems. Notably, MARDU worst-case overhead is significantly better than similar systems. The average overhead of MARDU is 5.5%, and the worst-case overhead is 18.3% (perlbench); in comparison to Shuffler [88] and CodeArmor [29], whose reported average overheads are 14.9% and 3.2%, respectively, while their worst-case overhead are 45% and 55%, respectively (see Table 1). This confirms MARDU is capable of matching if not slightly.

| Benchmark | Return address | ROP gadget | Func reuse | MARDU reduction (%) |
|-----------|----------------|------------|-----------|---------------------|
| perlbench | 13963          | 272        | 34371     | 1668                |
| bzip2     | 277            | 56         | 1569      | 81                  |
| g++       | 48096          | 518        | 89746     | 4318                |
| mcf        | 1173          | 77         | 313       | 33                  |
| mile       | 2104           | 7          | 3864      | 244                 |
| glogsh     | 9521           | 48         | 37999     | 2477                |
| hmuner     | 4237           | 12         | 9466      | 478                 |
| sjeng      | 1054           | 4          | 3110      | 139                 |
| libquantum | 469            | 3          | 1686      | 107                 |
| lddef      | 3118           | 372        | 14456     | 528                 |
| btm         | 70             | 3          | 394       | 26                  |
| sphinx3    | 2714           | 11         | 5628      | 526                 |
| perlbench  | 5316           | 309        | 15434     | 1565                |
| musl libc  | 4400           | 77         | 29743     | 3722                |

Total 95416 1695 247979 15712 345090/360802 (95.6%).
improving the performance (especially worst-case) overhead, while casting a wider net in terms of known attack coverage, compared to the current state-of-the-art.

**Performance overhead breakdown.** The two prominent sources of runtime overhead in MARDU are trampolines and the shadow stack. To understand how much runtime overhead each code transformation imposes, we ran SPEC with two different configurations: we first enabled only trampolines and do not use a shadow stack (trampoline only in Figure 6), then we enabled both trampolines and the shadow stack (full MARDU in Figure 6). The performance overhead of both are normalized to vanilla SPEC compiled with vanilla Clang. As Figure 6 shows, the major source of overhead is trampoline; trampolines incur 5.9% overhead on average, with the worst cases being 15.8% and 14.2% of overhead, for per1bench and h264ref, respectively. Note that MARDU’s shadow stack overhead is negligible. The average difference comparing MARDU trampolines with full MARDU is less than 0.3%, and in the noticeable gaps, adding less than 2% to MARDU in most cases compared to using MARDU trampolines only. This is expected as MARDU uses a compact shadow stack without out comparison epilogue. This implementation performs only essential bookkeeping to utilize the shadow stack; skipping unnecessary epilogue micro-optimizations [28].

### C. Scalability Evaluation

**Runtime performance overhead with NGINX.** NGINX is configured to accommodate a maximum of 1024 connections per processor, and its performance is observed according to the number of worker processes. wrk [43] is used to generate HTTP requests for benchmarking. wrk spawns the same number of threads as NGINX workers and each wrk thread sends a request for a 6745-byte static html. To see the worst-case performance, wrk is run on the same machine as NGINX to factor out network latency. Figure 7 presents the performance of NGINX with and without MARDU for a varying number of worker processes. The performance observed shows that MARDU exhibits very similar throughput to vanilla. MARDU incurs 4.4%, 4.8%, and 1.2% throughput degradation on average, at peak (12 threads), and at saturation (24 threads), respectively. Note that Shuffler [88] suffers from the overhead from per-process shuffling thread. Even in their NGINX experiments with network latency (i.e., running a benchmarking client on a different machine), Shuffler shows 15-55% slowdown. This verifies MARDU’s design that having the crashing process perform system-wide re-randomization, rather than a per-process background thread as in Shuffler, scales better.

**Load-time randomization overhead.** We categorize load-time to cold or warm load-time whether the in-kernel code cache (2 in Figure 3) hits or not. Upon a code cache miss (i.e., the executable is first loaded in a system), MARDU performs initial randomization including function-level permutation, start offset randomization of the code layout, and loading & patching of fixup metadata. As Figure 8 shows, all C SPEC benchmarks showed negligible overhead averaging 95.9 msec. gcc, being the worst-case, takes 771 msec; it requires the most overall fixups relative to other benchmarks (see Table III). For NGINX, we observe that load time is constant (61 msec) for any number of specified worker processes. Cold load-time is roughly linear to the number of trampolines in Table III. Upon a code cache hit, MARDU simply maps the already-randomized code to a user-process’s virtual address space. Therefore we found that warm load-time is negligible. Note that, for a cold load-time of musl 11bc takes about 52 msec on average. Even so, this is a one time cost; all subsequent warm load-time accesses of fetching musl 11bc takes below 1 µsec, for any program needing it. Thus, load time can be largely ignored.

**Re-randomization latency.** Figure 9 presents the time taken to re-randomize all associated binaries of a crashing process. The time includes creating re-randomizing the code layout, and reclaiming the old code (1 4 in Figure 4). To emulate an XoM violation, we killed the process with a SIGBUS signal and measured the re-randomization time inside the kernel. The average latency of SPEC is 6.2 msec. The difference between load-time and re-randomization latency because MARDU takes advantage of the metadata being cached from load-time, this means no redundant file I/O penalty is incurred, giving this performance gain. To evaluate the efficiency of re-randomization
on multi-thread/multi-process applications, we measured the re-randomization latency with varying number of NGINX worker processes up to 24. We confirm that the latency is consistent regardless of number of workers (5.8 msec on average with 0.5 msec of standard deviation).

**Re-randomization overhead under active attacks.** A good re-randomization system should exhibit good performance not only in its idle state but also under stress from active attacks. To evaluate this, we stress test MARDU under frequent re-randomization to see how well it can perform, assuming a scenario that MARDU is under attack. In particular, we measure the performance of SPEC benchmarks while triggering frequent re-randomization. We emulate the attack by running a background application, which continuously crashes at the given periods: 1 sec, 100 msec, 50 msec, 10 msec, and 1 msec. SPEC benchmarks and the crashing application are linked with the MARDU version of musl libc, forcing MARDU to constantly re-randomize musl libc and potentially incur performance degradation on other processes using the same shared library. In this experiment, we choose three representative benchmarks, milc, sjeng, and gobmk, that MARDU exhibits a small, medium, and large overhead in an idle state, respectively. Figure 10 shows that the overhead is consistent, and in fact, is very close to the performance overhead in the idle state observed in Figure 5. More specifically, all three benchmarks differ by less than 0.4% at a 1 sec re-randomization interval. When we decrease the re-randomization period to 10 msec and 1 msec, the overhead is quickly saturated. Even at 1 msec re-randomization frequency, the additional overhead is under 6%. These results confirm that MARDU provides performant system-wide re-randomization even under active attack.

**Runtime memory savings.** While an upfront one-time cost is paid for instrumenting with MARDU, the savings greatly outweigh this. To illustrate, we show a typical use case of MARDU in regards to shared code. musl libc is ≈800 KB in size, instrumented is 2 MB. Specifically, musl libc has 14K trampolines and 7.6K fixups for PC-relative addressing, the total trampoline size is 190 KB and the amount of loaded metadata is 1.2 MB (refer to Table III). Since MARDU supports code sharing, only one copy of libc is needed for the entire system. Our experimental setup at idle reported 310 processes while a typical 2-core consumer laptop at idle reported 263 processes. Therefore a rough estimate of memory savings for libc that MARDU provides compared to similar instrumentation that did not support code sharing is over ≈526-620 MB of still usable runtime memory. Furthermore, comparing to time-based continuous re-randomization techniques such as Shuffler [88] and CodeArmor [29] which almost always maintain two copies of code, MARDU’s memory saving for libc is ≈1-1.2 GB. Backes et al. [22] and Ward et al. [85] also highlighted the code sharing problem in randomization techniques and reported a similar amount of memory savings by sharing randomized code. Finally, note that the use of shadow stack does not increase runtime memory footprint because MARDU solely relocates return address from the normal stack to the shadow stack.

**VII. DISCUSSION**

**Applying MARDU to binary programs.** Although MARDU requires access to source code, applying MARDU directly to binary programs is possible. The job of the MARDU compiler is to detect all indirect control transfers (call/ret) and instrument such transfers to utilize trampolines. Such instrumentation can be done at binary-level if each call/ret can be precisely detected in a program. Applying MARDU to non-PC relative binary is challenging; however, position-independent executables (PIE), are now the default code generation mode in gcc-7, which naturally use PC-relative
addressing. Therefore, MARDU should be practical enough to adopt with little additional effort.

**Full-function reuse attacks.** Throughout our analysis, we show that existing re-randomization techniques that use a function trampoline or indirection table, i.e., use immutable (indirect) code pointer across re-randomization, cannot prevent full-function reuse attacks. This also affects MARDU; although limited to functions exposed in the trampoline, MARDU cannot defend against an attacker reusing such exposed functions as gadgets by leaking code pointers. We believe that this is a limitation of using immutable code pointers, and one possible solution to prevent these attacks could be pairing MARDU together with control-flow-integrity (CFI) [15], [30], [40], [45], [46], [60], [64], [65], [68], [71], [80], [81], [83], [89], [90] or code-pointer integrity/separation (CPI/CPS) [58]. MARDU already provides backward-edge CFI via shadow stack, so forward-edge CFI can also be leveraged to further reduce available code-reuse targets. MARDU’s defense is orthogonal to CFI, so applying both defenses can complement each other to provide better security. However, completely eliminating full-function code re-use and data-oriented programming [49] with low performance overhead and system-wide scalability is still an open problem.

**VIII. CONCLUSION**

While current defense techniques are capable of warding off current known ROP attacks, most designs must inherently tradeoff well-rounded performance and scalability for their security guarantees. With this insight, we introduce MARDU, a novel on-demand system-wide re-randomization technique to combat code-reuse attacks. MARDU shows pragmatic defense design is indeed navigatable in the face of the jungle that the code reuse attack landscape is. MARDU is the first code-reuse defense capable of code-sharing with re-randomization and thus allows scalability in an effort to focus on practicality. In addition, by being able to re-randomize on-demand, MARDU eliminates both the costly runtime overhead and the integral component of a threshold associated with continuous re-randomization. Our evaluation verifies MARDU’s security guarantees against known attacker models and adequately quantifies its high-level of entropy. Furthermore, MARDU’s performance evaluation showcases its robustness when deployed with real-world applications derived from SPEC CPU2006 averaging overhead of 5.5% as well as confirming scalability in multi-process scenarios with NGINX web server, averaging only 4.4% degradation.

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