GARMR: Defending the gates of PKU-based sandboxing

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Abstract
Memory Protection Keys for Userspace (PKU) is a recent hardware feature that allows programs to assign virtual memory pages to protection domains, and to change domain access permissions using inexpensive, unprivileged instructions. Several in-process memory isolation approaches leverage this feature to prevent untrusted code from accessing sensitive program state and data. Typically, PKU-based isolation schemes need to be used in conjunction with mitigations such as CFI because untrusted code, when compromised, can otherwise bypass the PKU access permissions using unprivileged instructions or operating system APIs.

Recently, researchers proposed fully self-contained PKU-based isolation schemes that do not rely on other mitigations. These systems use exploit-proof call gates to transfer control between trusted and untrusted code, as well as a sandbox that prevents tampering with the PKU infrastructure from untrusted code.

In this paper, we show that these solutions are not complete. We first develop two proof-of-concept attacks against a state-of-the-art PKU-based memory isolation scheme. We then present GARMR, a PKU-based sandboxing framework that can overcome limitations of existing sandboxes. We apply GARMR to several memory isolation schemes and show that it is practical, efficient and secure.

1 Introduction
Many computer programs contain untrusted components that must be isolated from trusted components to guarantee the confidentiality and integrity of sensitive program state or data. Modern operating systems provide the necessary isolation only at the process boundary. This forces software developers to run components as separate processes (sometimes referred to as compartments) that each have their own virtual address space. One of the drawbacks of process-level compartmentalization is that synchronous interaction between isolated components incurs high performance overhead due to the expensive context-switching required for inter-process communication. The research community has, therefore, proposed several alternative forms of compartmentalization that have better performance characteristics and are often more practical to apply to existing programs. The core idea behind many of these techniques is to isolate untrusted components in-process, thereby allowing them to share typical per-process resources with the rest of the program [4, 9, 11, 24, 25, 27, 28, 31, 36, 39, 42, 43, 45, 46, 49, 50, 56].

Memory Protection Keys for Userspace (PKU) is a hardware feature that is available on recent Intel and AMD server and desktop CPUs [13]. PKU allows programs to be compartmentalized by assigning memory pages to memory protection domains whose access permissions can be set individually by modifying the content of the PKU control register (PKRU). PKU exposes a set of unprivileged instructions to read and modify said register. This allows a program to quickly disable access to all compartments that must be isolated from the currently executing compartment, without having to pay the high cost of the system calls and TLB invalidations required to change page permissions through conventional means. However, an attacker can exploit vulnerabilities, hijack the control-flow of a program and abuse PKRU-updating instructions to modify PKRU, enabling access to any compartment’s memory.

Recently, researchers proposed ERIM [50] and Hodor [27], two efficient PKU-based memory isolation schemes that isolate an application’s trusted from its untrusted components by placing them in different memory protection domains. Both systems have built-in sandboxes that prevent adversaries from bypassing the isolation scheme by executing unsafe instructions that modify the PKRU register. Similar to previous work, we refer to these sandboxes as PKU-based sandboxes [12]. Unfortunately, these sandbox implementations have known security and usability problems. ERIM, for example, relies on static binary instrumentation (SBI) to neutralize any unsafe instructions in the protected program. However, as SBI cannot reliably distinguish code from data, ERIM could leave some unsafe instructions untouched [57]. Currently, ERIM’s sandbox also marks pages that contain
unsafe instructions as non-executable, which could lead to usability issues [50]. Hodor’s sandbox uses hardware breakpoints to ensure the program cannot execute unsafe instructions. This approach does not rely on SBI like ERIM’s, but both systems can still be bypassed using the kernel as a confused deputy [12].

In addition to the aforementioned problems, we found two more security flaws in the design of Hodor’s sandbox and managed to exploit these flaws in new proof-of-concept attacks that we present in this paper. We then designed and implemented GARMR, a new PKU-based sandboxing framework that can protect PKU-based memory isolation schemes against all known PKRU-tampering attacks (except for the signal context attacks described in Section 4.2.7). We used GARMR to develop sandboxes for two existing isolation schemes and evaluated the performance and security of the resulting systems. We conclude that GARMR enables practical, efficient, and secure PKU-based sandboxing.

In summary, our paper contributes the following:

- We identified new design flaws in Hodor’s PKU-based sandbox and developed two new proof-of-concept attacks that exploit these flaws [27].
- We present GARMR, a new PKU-based sandboxing framework, and apply our framework to develop sandboxes for two state-of-the-art PKU-based memory isolation schemes: ERIM [50] and XOM-Switch [39]. The resulting sandboxes stop all known attacks (except for the signal context attacks described in Section 4.2.7), including the new attacks we present in this paper [12, 27, 50].
- We perform an extensive evaluation of the constructed sandboxes on real-world server applications and show that GARMR enables practical, efficient, and secure PKU-based sandboxing.

2 Background

PKU utilizes a new user-mode register (PKRU) to control access rights\(^1\) to memory pages that are tagged with one of 16 available protection keys. The PKRU register is 32 bits wide and has two bits (access disable and write disable) for each key. These bits are checked during memory accesses for all the pages that are associated with a key. The OS provides new system calls, pkey_alloc and pkey_free, to allocate and free protection keys respectively. A process can tag a page with a key by using the new pkey_mprotect system call and access the PKRU register with unprivileged instructions; rdrpkr for read and wrpkr for write accesses. The xrstor instruction can also update the PKRU register if bit 9 in the eax register is set prior to the instruction execution.

\(^1\)The PKRU register only controls data accesses, instruction fetches are not similarly restricted.

2.1 PKU-based Memory Isolation Schemes

Some memory isolation schemes leverage PKU to isolate trusted from untrusted components [24, 27, 50]. These systems typically tag memory pages containing trusted code and data with a different protection key than pages containing untrusted code, thereby placing them in different memory protection domains. Trusted code is then allowed to access both domains, whereas untrusted code can only access the untrusted domain. Researchers have also used PKU to harden JavaScript engines [43], reinforce other exploit mitigations [9, 25, 31], and provide software abstractions for isolation and sandboxing [42].

One major challenge for PKU-based systems is to prevent attackers from tampering with the PKRU register by exploiting a memory vulnerability in the untrusted code and by subsequently locating and executing a PKRU-modifying instruction. One way to prevent such attacks is to apply exploit mitigations to the untrusted code [2, 8, 33]. However, these mitigations can introduce non-trivial run-time overheads and often cannot fully prevent PKRU tampering.

Hodor and ERIM are PKU-based isolation systems that do not rely on such mitigations [27, 50]. Both approaches use two domains, M\(_T\) and M\(_U\), that contain trusted and untrusted components respectively. Transferring control from trusted components to untrusted components or vice versa happens via well-defined, exploit-proof instruction sequences also known as call gates. Hodor and ERIM each use a sandbox to prevent attackers that manage to compromise an untrusted component from accessing M\(_T\) by abusing PKRU-modifying instructions (wrpkr, xrstor) or system calls like pkey_mprotect as depicted in Figure 1.
2.2 A Closer Look at ERIM

ERIM compartmentalizes applications through binary rewriting. It assumes the trusted components T are not exploitable, but does not make any assumptions about the untrusted components U. ERIM implements call gates using so-called safe instructions. Safe wrpkrus instructions are those that are immediately followed by either instructions to validate PKRU’s state at run time (ensuring that M_T is locked by PKRU) or by a jump to T. Safe xrstors instructions, on the other hand, are immediately followed by instructions that check if bit 9 of eax is set. If one of these run-time validations fails, the control flow jumps to an instruction sequence that terminates the application. Otherwise, the program execution continues. We refer to any other wrpkrus and xrstors instructions as unsafe instructions similar to previous work [12]. ERIM’s call gates are not exploitable, since they do not contain unsafe instructions. However, an attacker that controls U could abuse unsafe instructions found outside call gates to change PKRU, thereby allowing U to access M_T. Attackers can easily find unsafe instructions because (a) they could appear as operands of other instructions, and (b) x86 instructions do not have a fixed size and the CPU, therefore, allows programs to execute instruction operands as if they were regular instructions themselves.

At startup time, ERIM’s PKU-based sandbox scans all executable pages of the protected application using the /proc/self/mem interface to verify the absence of exploitable unsafe instructions in M_T pages. Any executable page that contains unsafe instructions is marked as non-executable. Consequently, any attempt to execute code from such a page will trigger a fault that is handled by ERIM’s sandbox, and the sandbox terminates the program prematurely. To prevent this, ERIM first uses a static binary rewriter to replace instruction sequences that contain unsafe wrpkrus and xrstors by functionally equivalent sequences that do not contain such unsafe occurrences. At run time, the sandbox intercepts and monitors mmap, mprotect and pkey_mprotect system calls from U that can introduce unsafe instructions or allow access to M_T. ERIM provides two sandbox implementations: one version that is based on ptrace [1], and a more efficient one that requires minor kernel modifications.

2.3 A Closer Look at Hodor

Hodor partitions applications into trusted and untrusted libraries. Hodor’s trusted application loader ensures that untrusted libraries U can only interact with trusted libraries T through call gates similar to ERIM’s. However, unlike ERIM, Hodor does not rely on code rewriting.

Hodor’s PKU-based sandbox monitors the application at run time to stop an attacker from abusing unsafe wrpkrus instructions that change PKRU. When an application attempts to mark a page as executable, the trusted loader first scans the page for unsafe wrpkrus instructions and marks the page as non-executable if it contains unsafe instructions. Attempts to execute code from such a page trigger a fault. Upon receiving that fault, Hodor’s modified OS kernel attempts to put hardware breakpoints on all unsafe instructions on the same page, and it marks the page as executable. If the page contains more unsafe wrpkrus instructions than the maximum number of hardware breakpoints the CPU supports, Hodor will single-step through the page instead [1]. This mechanism ensures that all unsafe wrpkrus instructions will be vetted by Hodor’s kernel. When a hardware breakpoint is triggered, Hodor’s modified kernel terminates the program execution. Hodor can reclaim hardware breakpoint slots in the debug registers if and when necessary. However, when doing so, it will always mark the pages these breakpoints point to as non-executable.

2.4 XOM-Switch

PKU can also be used for eXecute-Only Memory (XOM). Recent kernels (Linux 4.9+) support XOM through an enhanced version of the mprotect system call2. Support for XOM is missing in libc and compilers, though. XOM-Switch [39] patches the dynamic loader/linker and libc to apply XOM memory to ELF binaries. However, the defense is limited in power unless it is combined with an additional mitigation such as CFI, since an attacker can abuse PKU to disable XOM. The authors of XOM-Switch acknowledged its limitations and proposed the use of Intel’s Control-flow Enforcement Technology (CET) [1], an upcoming feature of Intel CPUs, as the ultimate solution against such attackers.

3 Threat Model

For this paper, we make the following assumptions about the host system, the targeted application and the attacker. Our assumptions are in line with work in the area [12, 27, 50]:

- **Host System.** We assume Protection Keys for Userspace (PKU) [13] to be available on the target platform and we trust its implementation. The kernel is also considered part of the Trusted Computing Base (TCB).

- **Targeted Application.** We do not make any assumptions about the untrusted code U, but similar to previous work [12, 27, 50], we assume that the initial state of the targeted application is not compromised and that the PKU-based sandbox is initialized correctly. The trusted code T of the application, however, is considered free of exploitable bugs. We also assume for simplicity that there are only two levels of trust (T and U), and that the application is not using PKU for any purposes, except for memory isolation.

- **Attacker.** We consider an attacker that controls U of the targeted application with the goal to access

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2The kernel assigns keys to memory pages and updates the PKRU register accordingly. No user space code is involved.
M_T. For example, an adversary can use code-reuse and control-flow hijacking attacks to exploit unsafe instructions in executable pages of M_T to tamper with PKRU’s state, enabling access to M_T. Mitigations like software diversity [33] and CFI [2, 8] raise the bar for such attacks, but we do not rely on such defenses. Attacks that target the underlying hardware such as transient execution [23, 30, 35] and remote-fault injection [5, 29, 47, 51] are considered out of scope for this paper.

4 Challenges
ERIM and Hodor have built-in sandboxes to prevent attackers that control U from accessing M_T by any means. However, there are several security and usability challenges associated with PKU-based sandboxing [12, 27, 50]. We describe these issues below, introduce two new attacks against Hodor, and also discuss potential solutions.

4.1 Handling of unsafe instructions
ERIM and Hodor eliminate or detect unsafe instructions that can tamper with PKRU’s state. We show below that the proposed techniques to handle unsafe instructions lead to security and usability issues.

First, Hodor’s sandbox ensures that unsafe wrpkru instructions are vetted by its modified kernel, leveraging hardware breakpoints and single-step execution. We carefully examined the open source implementation of Hodor3 and discovered that it does not monitor unsafe xrstor instructions. Therefore, an attacker that controls U can abuse these unsafe xrstor instructions to unlock M_T.

ERIM, on the other hand, relies on SBI to neutralize unsafe wrpkru and xstor instructions. In addition, ERIM’s sandbox inspects the program at run time to ensure no new unsafe instructions are introduced in the executable pages of M_T. If U tries to map a page that contains unsafe instructions, it is marked as non-executable by ERIM’s sandbox to stop attackers from exploiting them. We used the open source implementation of ERIM4 and repeated the experiments described in the original paper [50]. We verified that ERIM’s approach works for the tested system (Debian 8, Linux 4.9.60) and applications.

However, we could not replicate the experiments on recent systems (Ubuntu 18.04, Linux 5.3.18). The reason is that ERIM’s tested system is fairly old, and uses outdated versions of binaries such as the dynamic linker ld.so, libc.so, and libm.so (version 2.19). ERIM’s rewriting strategy cannot eliminate all unsafe instructions in recent versions of these binaries (version 2.27), either due to SBI’s inability to distinguish code from data [57] or limitations of the current prototype. Furthermore, ERIM’s sandbox marks pages containing unsafe instructions as non-executable, creating usability issues as described in Section 1 of ERIM [50]. Even trivial compartmentalized programs are terminated early by ERIM’s sandbox in recent systems, since they attempt to execute code in pages that contain unsafe instructions. We can modify the sandbox to not mark these pages as non-executable, but this would lead to security issues, since an attacker that controls U can exploit unsafe instructions from these pages.

4.2 PKU Pitfalls
Conor et al. [12] developed proof-of-concept exploits against ERIM and Hodor that bypass their PKU-based sandboxes. These attacks use the kernel as a confused deputy, taking advantage of OS abstractions that are agnostic of PKU-based memory isolation schemes. We briefly describe them below.

4.2.1 Inconsistencies of PT and PKU permissions.
The OS exposes system calls that do not respect the enforced page table (PT) and PKU permissions. An attacker can use process_vm_readv, process_vm_writev and ptrace system calls to directly access M_T from U. PKU-based sandboxes should intercept and monitor these system calls to prevent such accesses. This introduces negligible performance overhead, since these calls are rarely used, as shown in previous work [12].

Another method to circumvent enforced page table and PKU permissions is to use the procfs interface. A process can open its /proc/self/mem file, in which positions correspond to addresses in the process’ address space, and perform I/O operations on it. By design these operations ignore page table and PKU permissions, allowing an attacker to directly access M_T from U, or modify non-writable code to tamper with PKRU’s state. PKU-based sandboxes should at least intercept and monitor open-like system calls to prevent such attacks. However, it is shown that this adds huge overhead, unless an efficient system call interception and monitoring mechanism is used [12].

4.2.2 Mappings with mutable backings.
More problems arise when mapped memory is backed by a mutable file. An attacker can directly perform I/O operations on this file and modify it, regardless of the page table and PKU permissions of the mappings that are backed by it. These modifications are reflected to the corresponding mappings. The OS also allows multiple mappings of the same shared memory with different page permissions that refer to the same physical memory. Therefore, an attacker can modify an immutable and executable mapping through another writable mapping of the same shared memory. In both mentioned cases, the attacker can add unsafe instructions to executable pages without being detected by the PKU-based sandbox.

4.2.3 Changing code by relocation.
Attackers can also introduce unsafe instructions, without modifying executable pages of \( M_T \). First, they can create two non-neighboring mappings that each contains part of an unsafe instruction at the page boundary. The PKU-based sandbox scans the mappings for unsafe instructions, and both are considered safe, since they do not contain complete unsafe instructions. Then attackers can use the \texttt{mremap} system call to move the mappings next to each other, and form an unsafe instruction. To stop this attack, the sandbox should re-scan the page boundaries for unsafe instructions after every relocation [12].

4.2.4 Influencing intra-process behavior with seccomp.

ERIM and Hodor use the new \texttt{pkey_mprotect} system call to isolate \( M_T \) from \( U \). The trusted code \( T \) allocates a dedicated memory region \( (M_T) \) to store secrets, e.g., encryption keys, and uses \texttt{pkey_mprotect} to associate it with a different protection domain than \( U \). However, a malicious seccomp filter can deny these calls and return a success value, tricking \( T \) to store sensitive data in memory that is not properly isolated from \( U \). A sandbox can prevent this attack by intercepting and restricting \texttt{prctl} and seccomp system calls [12].

4.2.5 Modifying trusted mappings.

Attackers can also change the virtual address space to access isolated memory or modify \( T \). For example, an adversary can invoke a \texttt{pkey_mprotect} system call to modify the protection key that a page of \( M_T \) is tagged with. Hodor prevents such attacks by passing the addresses of \( T \) and \( M_T \) to the kernel, and denying any attempt to change them from \( U \) [27].

4.2.6 Race conditions.

PKU-based sandboxes scan executable pages for unsafe instructions. An attacker who controls multiple threads can exploit race conditions in the memory scanning process to bypass these sandboxes and add unsafe instructions to executable memory. To protect against such attacks, the sandbox should initially mark the page as non-writable and non-executable, scan it, and then mark it as executable in case that it does not contain unsafe instructions [12]. Furthermore, ERIM’s sandbox cannot reliably determine trusted mappings (requested by \( T \)), allowing an attacker to trick the sandbox to accept untrusted mappings (originated from \( U \)) as trusted ones.

4.2.7 Signal context attacks.

An adversary can abuse signals to modify PKRU’s state. The CPU state is exposed both during signal delivery, and the return from a signal handler. An attacker can use system calls such as \texttt{sigreturn} and \texttt{signalstack} to tamper with PKRU’s state directly or indirectly without using \texttt{wrpkru} or \texttt{xrstor} instructions. For example, an attacker can craft a CPU state on the stack and use the \texttt{sigreturn} system call to restore an arbitrary value to the PKRU register [12].

4.3 New PKU Pitfalls

We examined the open source implementation of Hodor, and we developed two new proof-of-concept attacks against it that bypass the memory isolation scheme and are not stopped by its sandbox. Our attacks, specifically target Hodor’s instruction vetting mechanism that uses debug registers to monitor unsafe instructions.

4.3.1 Vetted unsafe instruction relocation.

Hodor’s modified kernel loads the addresses of unsafe instructions of an executable page in debug registers, and terminates execution if a hardware breakpoint is triggered. An attacker can relocate this page using \texttt{mremap} system call. Hodor’s sandbox, however, does not intercept this call. Consequently, debug registers are not updated, allowing exploitation of unsafe instructions in that specific page. This is a different attack than the one described in Section 4.2.3, and requires additional measures. To deal with this attack, Hodor’s sandbox should intercept \texttt{mremap} system call, mark the page as non-writable and non-executable, relocate the page, update the debug registers to store the new addresses of the unsafe instructions, and finally mark the relocated page as non-writable and executable.

4.3.2 Incomplete debug register update.

We also discovered a fundamental flaw in Hodor’s approach to monitor unsafe instructions. The key problem is that threads share code and data, while debug registers are unique to each thread. Hodor’s sandbox vets unsafe instructions using debug registers as described in Section 2.3. However, when one thread maps memory containing unsafe instructions, it is not possible to update debug registers for all threads synchronously, unless a process-wide stop the world mechanism is implemented. As a result, an attacker can use one thread to introduce unsafe instructions, and abuse them from another thread to tamper with PKRU’s state and unlock \( M_T \) to \( U \).

5 Design, Workflow and Use Cases

5.1 Design

With the aforementioned challenges in mind, we designed and implemented GARMR, a general framework that developers can use to build PKU-based sandboxes. GARMR’s design draws inspiration from the following three observations:

1. Instruction vetting techniques such as the ones used in Hodor do not need to be implemented in the kernel. Instead, we can implement them in a user-space tool using the \texttt{ptrace} API.
2. Static binary rewriting techniques such as those used in ERIM can eliminate many, but not necessarily all unsafe instructions in an application (see Section 4.1).

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\[^{5}\text{This attack targets the ptrace-based implementation of ERIM’s sandbox.}\]
Figure 2. The developer (optionally) uses the SBI tool to eliminate as many unsafe instructions as possible from the applications. We represent the eliminated unsafe instructions with strike-through text in the rewritten application. Next, the developer uses GARMR APIs to extend the GARMR Loader and GARMR Monitor. The produced PKU-based sandbox consists of the modified loader, the modified monitor and the syscall agent.

Eliminating unsafe instructions is useful even for sandboxes that use hardware breakpoints for instruction vetting, because doing so reduces pressure on the processor’s debug registers.

3. PKU-based sandboxes should intercept and monitor only a few system calls, and the vast majority of these calls are not frequently invoked by applications. System calls used to open file descriptors (e.g., sys_open) are an exception.

Our framework consists of an SBI tool (Binary Rewriter), a ptrace-based monitor (GARMR Monitor), a custom loader (GARMR Loader), a set of APIs to extend the loader and the monitor (GARMR APIs), and a minimal in-kernel syscall agent (Syscall Agent). We describe these components below.

**Binary Rewriter.** SBI cannot neutralize all unsafe instructions (see Section 4.1). However, an SBI tool can still eliminate a portion of unsafe instructions from the binaries. Consequently, the PKU-based sandbox has to monitor fewer unsafe instructions. Currently, we use ERIM’s SBI tool to rewrite binaries [50].

**GARMR Loader.** We developed a custom application loader that, upon startup, maps a special executable and non-writable page containing an rdpkru instruction into the application’s address space. We use this instruction to implement an interface for a ptrace-based monitor to read PKRU register (see GARMR APIs below). Then, the loader initializes the syscall agent through a prctl option that we added to the kernel, and completes the program loading.

**GARMR Monitor.** We provide a ptrace-based monitor that intercepts and monitors system calls. The monitor also injects the custom loader into the application before launching it, and ensures that the access permissions of the aforementioned special page are not changed and that it is not unmapped. The monitor can (optionally) log the invoked system calls and their results. GARMR spawns a new monitor for every newly created process or thread.

**Syscall Agent.** ptrace-based monitors introduce high overhead due to expensive context-switching and TLB invalidations. To avoid this cost, we implemented a minimal in-kernel syscall agent that can let certain system calls bypass the ptracer. Each thread has its own agent. Our custom loader initializes the agent, providing a list of system calls (SList) and a list of inodes (IList). The agent forwards to GARMR monitor only the system calls of SList, while the rest are executed natively, bypassing ptrace. It also denies opening of sensitive files included in IList. This functionality can also be implemented in a ptrace-based monitor like the GARMR monitor, but it would add substantial run-time overhead, as shown in previous work [12].
GARMR APIs. GARMR implements generic macros and functions that developers can use to extend the GARMR monitor and the GARMR loader to implement PKU-based sandboxes, tailored to the needs of the memory isolation scheme. We briefly describe the most important APIs. First, developers can use simple macros to add/remove items to/from SLlist and IList. In addition, GARMR provides macros and functions to write system call handlers, and to access registers and memory. Developers can also use APIs to scan the application’s memory for unsafe instructions on executable pages, similar to ERIM and Hodor.

Moreover, our framework provides a ptrace-based interface to read PKRU, reliably identifying the current domain permissions. Specifically, we use ptrace to change control flow and execute the rdpkru instruction that the loader mapped into the address space at startup. This functionality is important, since the current version of the ptrace API cannot access the PKRU. Finally, GARMR exposes APIs to access debug registers and to set/unset hardware breakpoints, an emulation engine for x86 instructions and a rich logging infrastructure that can help sandbox developers identify and fix bugs. We describe use cases for these APIs in Section 5.3.

5.2 Development Workflow

The first (optional) step of the development workflow is to eliminate a portion of the unsafe instructions that are present in the to-be-sandboxed application with an SBI tool. We rewrite the application binary, system libraries and other dependencies. Then, the developer uses the provided APIs to extend the loader and the monitor to build PKU-based sandboxes for different memory isolation schemes. For example, the developer has the freedom to implement different system call policies, define which instructions are safe and which are not, and use debug registers to vet unsafe instructions.

Each produced PKU-based sandbox consists of the following components: a modified loader, a modified monitor and the syscall agent. For simplicity, we refer to the first two components as the loader and the monitor respectively in the rest of the paper. Note that GARMR does not modify the syscall agent. The same agent implementation is used across different sandboxes created with GARMR. Our framework only provides interfaces to change how the loader and the monitor interact with the agent. The full sandbox development workflow is shown in Figure 2.

5.3 Use Cases

We apply the aforementioned development workflow to build PKU-based sandboxes for ERIM and XOM-Switch. We describe the sandbox’s components (the monitor, the loader and the syscall agent) below and their interactions for each use case.

5.3.1 Use Case 1 – A Sandbox for ERIM.

The PKU-based sandbox intercepts and monitors the following system calls: modify_lutd, prctl, seccomp, ptrace, process_vm_readv, process_vm_writev, mprotect, pkey_mprotect, pkey_alloc, pkey_free, mmap, munmap, mremap, execve, shmat, shmdt, all variants of the fork/clone system calls, and all variants of the open system call. The loader adds these system calls to the SLlist such that our system call agent reports them to the monitor. We exclude open-like system calls from SLlist, however, because many applications invoke these system calls so frequently that monitoring them from user space would add substantial run-time overhead [12]. The loader also adds the inodes of all sensitive files to IList. Next, the loader passes these two lists to the syscall agent and initializes it.

At run time, the agent forwards the system calls included in SLlist to the monitor, while the rest are executed natively, bypassing ptrace. It also intercepts open-like calls directly in kernel space and denies opening of sensitive files to protect against attackers that aim to abuse these files to bypass memory isolation (see Section 4.2.1). Furthermore, the agent blocks opening of hard and soft links to sensitive files, since they use the same inodes. The monitor also ensures that the agent is only initialized once after execve or fork/clone-like system calls.

The monitor can inspect the PKRU register to reliably determine whether the program is currently executing code in U or T. This is necessary to stop certain attacks that rely on data races to bypass the sandbox (see Section 4.2.6). In addition, it tracks pages which are in M_T by intercepting pkey_mprotect system calls from T. An attacker that controls U can use system calls to access M_T directly or to unlock it to U (see Sections 4.2.1, 4.2.4 and 4.2.5). To prevent such attacks, our monitor forbids the following system calls from U: modify_lutd, prctl setting seccomp, seccomp, ptrace, pkey_alloc, pkey_free, pkey_mprotect, shmat and shmdt. It also rejects process_vm_readv and process_vm_writev system calls that attempt to access M_T from U.

The monitor scans all executable pages U loads into the address space at run time for unsafe instructions. To do so, it intercepts and monitors mmap, mremap and mprotect system calls from U as these calls could introduce unsafe instructions or unlock M_T. The memory scanning process ignores our call gates and xрастor instructions that are followed by proper checks. These sequences are also considered safe by ERIM. We implemented a modified version of Hodor’s instruction vetting scheme in user space using ptrace to deal with the discovered unsafe instructions.

Unlike Hodor, however, our scheme vets both unsafe wrpkr u and xрастor instructions. In addition, Hodor’s sandbox

\[\text{Currently, IList only contains the inode of } /proc/self/mem. However, we can easily expand IList in case that we discover more sensitive files.}\]
always terminates execution whenever a hardware break-
point is triggered, while our monitor allows execution of
unsafe \texttt{xrst}or instructions to continue if bit 9 of eax register
is not set. If the bit is set, we terminate the program. The mon-
tor also implements techniques described in Sections 4.2.3
and 4.3.1 to deal with attacks that use code relocation to
introduce unsafe instructions. Eliminating unsafe instruc-
tions using binary rewriting reduces debug register pressure
and, thus, decreases the likelihood that we have to resort
to single-step execution when a page contains more unsafe
instructions than the number of debug registers (see Sec-
tion 2.3).

Similar to previous work \cite{12}, our monitor enforces
W^X by intercepting and monitoring system calls that map pages
and change permissions, since Linux by default allows pages
that are both writable and executable. Moreover, the monitor
does not allow executable mappings that are MAP\_SHARED or
MAP\_SHARED\_VALIDATE to protect against attackers that try
to modify an immutable mapping via another mutable map-
ning of the same shared memory (see Section 4.2.2). Dealing
with attacks that try to directly modify the underlying object
of a file-backed mapping is more complicated, though. First,
the monitor intercepts system calls that map memory (e.g.,
mmap) and replaces the file-backed mappings, with MAP\_-
ANONYMOUS ones, ensuring there are no mappings that are
backed by a file. Then the monitor copies the file contents
that the application initially attempted to map to the mapped
region. To prevent attacks through memory pages with mu-
table backing files (see Section 4.2.2), the monitor needs to
impose restrictions on mapped regions, e.g., we reject any
attempt to map pages that are simultaneously MAP\_SHARED
and executable. Although we did not observe any compati-
bility or usability issues that arose from these restrictions,
we discuss their potential implications in Section 8.

\textbf{Protecting Multi-Threaded Programs.} Protecting
multi-threaded programs is a challenge because hardware
breakpoints are only set in the thread-local register context
and because malicious threads could attempt to modify
executable code while it is being scanned for unsafe
instructions. This opens new possibilities for attacks that
would not be possible in single-threaded programs, as we
explain in Sections 4.2.6 and 4.3.2. To counter these threats,
we spawn a unique monitor thread for every application
thread we protect. Monitor threads that supervise threads
of the same process share data structures and \textit{always} enter
a critical section when they scan the application memory
for unsafe instructions. Monitor threads must enter the
same critical section when an application thread attempts
to execute a system call that could change the contents
of any of the pages that are being scanned (e.g., mremap).
This design avoids the problem of race conditions during
memory scanning.

Secondly, whenever a process spawns a thread for the first
time (i.e., when the program transitions from single-threaded
to multi-threaded execution), the monitors stop relying solely
on hardware breakpoints for instruction vetting, since up-
dates to the set of breakpoints would not propagate beyond
the currently executing thread as described in Section 4.3.2.
Specifically, the monitors only use hardware breakpoints to
vet the unsafe instructions whose addresses were stored in
the debug registers at the moment the process started to use
multiple threads. The monitor marks code pages containing
unsafe instructions, not already protected by debug regis-
ters, as non-executable. This includes code pages that were
mapped as non-executable by the monitor before switching
to multi-threaded execution, as well as new pages mapped
by the threads. Attempts to execute instructions from code
pages that were marked as non-executable by the monitor
trigger a fault. When the monitor is notified of faults on one
of these pages, which it can determine by inspecting the
instruction pointer, it does not terminate execution imme-
diately. First, the monitor checks if the instruction that was
about to get fetched is an unsafe instruction or not. If it is
an unsafe \texttt{wrprku} instruction, the monitor terminates execu-
tion, while if it is an unsafe \texttt{xrst}or, the monitor terminates
execution \textit{only} if bit 9 of eax is set. Otherwise, the instruc-
tion is considered safe and the monitor uses an emulation
engine for x86 instructions (included in GARMR) to emulate
the instruction, essentially updating the program’s state to
make it seem like the instruction was \textit{actually} executed.

At run time, the monitor can decide if emulation of in-
structions is necessary by detecting if multiple threads are
used. To do so, it intercepts system calls that create and de-
stroy threads, and checks the \texttt{/proc/self/task} file. Even
though emulating instructions in user space is slow, we
did not experience significant performance degradation in
our experiments that included multi-threaded applications,
since unsafe instructions are rare as described in previous
work \cite{27, 50}. In addition, eliminating a portion of unsafe
instructions with SBI decreases the number of instructions
that should be emulated, since the monitor needs to mark
fewer code pages as non-executable. We provide details for
the emulation engine in Section 6 and we discuss alternative
ways to deal with multi-threaded applications in Section 8.

Hodor implements memory isolation in a way similar to
ERIM. As a result, we can apply the sandbox we described
above to Hodor as well, requiring only minor modifications,
e.g., swapping call gates.

\textbf{5.3.2 Use Case 2 – A Sandbox for XOM-Switch.}
We also used GARMR to build a sandbox for XOM-Switch.
To the best of our knowledge, this is the first PKU-based
sandbox for XOM-Switch. This sandbox prevents attackers
from abusing PKU to disable XOM, and, thus, eliminates
the requirement of an additional defense like CFI to protect
XOM (see Section 2.4).
The XOM-Switch sandbox is similar to the sandbox described above. Therefore, we focus only on their implementation differences. First, we treat XOM-Switch as a PKU-based memory isolation scheme, in which U includes all the application code. As a result, the sandbox does not need to reliably determine the current domain (it is always U). Second, \( M_T \) consists of the eXecute Only Memory (XOM). Since XOM-Switch uses the enhanced version of \texttt{mprotect}, the monitor tracks which pages are in \( M_T \) by inspecting the arguments of \texttt{mprotect} calls from U that apply XOM. Third, the monitor considers all \texttt{wrpkr} instructions unsafe, since inter-domain transitions are performed by the kernel and XOM-Switch does not rely on user space \texttt{wrpkr} instructions for inter-domain transitions (see Section 2.4 for details). Last, the monitor denies any permission changes of XOM; if a page is marked as execute only, it remains like that for the entire execution.

### 6 Implementation Details

\textbf{GARMR} consists of the following components: an SBI tool, the \textbf{GARMR} monitor, the \textbf{GARMR} loader, the \textbf{GARMR} APIs, and syscall agent (see Figure 2). We used ERIM’s binary rewriter as the SBI tool. We implemented the \textbf{GARMR} monitor, \textbf{GARMR} loader and \textbf{GARMR} APIs in 9536 lines of code. The emulation engine is part of the provided APIs and currently supports 170 x86 instructions. We also implemented the syscall agent with a small kernel patch (79 lines of code) for Linux kernel 5.3.18. The \textbf{GARMR} monitor and \textbf{GARMR} loader are designed to be extensible. We extended them using the \textbf{GARMR} APIs to implement PKU-based sandboxes for ERIM and XOM-Switch. These extensions consist of 1007 lines of code. The sandboxes we constructed have only minor implementation differences, as mentioned in Section 5.3.2. We plan to open source our framework, evaluation scripts, and proof-of-concept attacks upon acceptance of this paper.

### 7 Evaluation

We evaluated the performance and the security of the PKU-based sandboxes created with GARMR.

#### 7.1 Performance

We ran our experiments on an HP Z6 G4 workstation with a 12-core Intel Xeon Silver 4214 CPU running at 2.20 GHz and 64 GB of RAM (Turbo-Boost and Hyper-Threading were disabled). The machine runs Ubuntu 18.04.6 LTS with version 5.3.18 of the Linux kernel. We applied a minimal kernel patch that implements the syscall agent. We evaluated the constructed PKU-based sandboxes on popular high performance server applications: nginx, lighttpd and redis. We ran a benchmarking client on a separate machine that is connected to the workstation through a gigabit Ethernet connection. The client machine has a 6-core Intel Core i7-8700K CPU running at 3.70 GHz and 64 GB of RAM (Turbo-Boost and Hyper-Threading were disabled). The client machine runs Ubuntu 18.04.6 LTS with version 5.4.0 of the Linux kernel. For nginx and lighttpd, we used wrk benchmark to request a 4KB page for 10 seconds over 10 concurrent connections. For redis, we used redis-benchmark, distributed with Redis, with the default workload (100000 requests and 50 parallel connections).

For the experiments described in Section 7.1.2, the client communicates with the server over HTTPS, while for the experiments described in Sections 7.1.1 and 7.1.3 communication happens over HTTP. We measured the throughput of the server applications running under our defense relative to the throughput of the native execution. We ran each experiment 10 times, removed the highest and lowest values as outliers, and reported the average of the 8 remaining values. We configured lighttpd and nginx to use 1–3 workers, and redis to use 1–3 I/O threads. The server applications can saturate the network connection when configured with 3

| APP                | ERIM-CPI (No Sandbox) | ERIM-CPI with GARMR Sandbox | ERIM-SS (No Sandbox) | ERIM-SS with GARMR Sandbox |
|--------------------|-----------------------|-----------------------------|----------------------|----------------------------|
| nginx (1 worker)   | 6.54%                 | 4.97%                        | 0.57%                | 1.91%                      |
| nginx (2 workers)  | 5.72%                 | 6.20%                        | 4.62%                | 3.32%                      |
| nginx (3 workers)  | 1.55%                 | 2.23%                        | -1.22%               | 1.05%                      |
| lighttpd (1 worker)| –                     | –                           | 0.16%                | 2.14%                      |
| lighttpd (2 workers)| –                    | –                           | 0.22%                | 0.23%                      |
| lighttpd (3 workers)| –                   | –                           | -0.01%               | -0.03%                     |
| geometric mean     | 3.87%                 | 4.10%                        | 0.71%                | 1.43%                      |

Table 1. We isolated shadow stacks and safe regions in CPI/CPS with ERIM (ERIM-SS and ERIM-CPI respectively). We measured the overhead of standalone ERIM-SS and ERIM-CPI when they are not protected by a sandbox (No Sandbox) compared to the native execution. We report the overhead of ERIM-SS and ERIM-CPI with GARMR sandbox compared to the native execution. GARMR sandbox here refers to the sandbox described in Section 5.3.1. – indicates that the experiment failed.
workers (lighttpd, nginx) and 3 I/O threads (redis). As a result, we did not try configurations with more than 3 workers and I/O threads.

The developed PKU-based sandboxes, described in Section 5.3, identified unsafe instructions in nginx, redis, ld.so, libm.so and libc.so during our experiments. We managed to eliminate an unsafe `xstor` instruction in nginx with ERIM's SBI tool, but we could not neutralize the other unsafe instructions with this tool. We did not investigate further the reason that the tool failed, since the constructed sandboxes do not solely rely on the SBI being successful in removing all unsafe instructions.

### 7.1.1 Protecting safe regions in CPI/CPS and SS.

Similar to previous work [50], we used ERIM to isolate safe regions of CPI/CPS [32]. We changed 13 lines of code to fix an LLVM bug and to port the CPI compiler of ERIM to our testing environment (Ubuntu 18.04 with kernel 5.3.18). In the same manner, we used ERIM to isolate the safe regions of a shadow stack implementation (SS for short) [9]. To do so, we added 38 lines of code to the SS compiler passes to add ERIM’s functionality. We refer to the above compiler passes as ERIM-CPI and ERIM-SS respectively. We applied ERIM-CPI to nginx and ERIM-SS to lighttpd and nginx. We could not run lighttpd after applying ERIM-CPI to it because of CPI’s imprecise handling of aliasing relations between memory references. We also verified that lighttpd fails with the original CPI compiler [32]. Similarly, we could not run redis after applying either ERIM-CPI or ERIM-SS. Again, we verified that redis also fails after compilation with the original CPI [32] and SS [9] compiler. Consequently, we concluded that it is not our code that is responsible for the failures.

| APP                  | ERIM-OpenSSL with GARMR Sandbox |
|----------------------|---------------------------------|
| nginx (1 worker)     | 1.12%                           |
| nginx (2 workers)    | 0.66%                           |
| nginx (3 workers)    | 1.21%                           |
| lighttpd (1 worker)  | 0.44%                           |
| lighttpd (2 workers) | 0.31%                           |
| lighttpd (3 workers) | 0.49%                           |
| redis (1 I/O thread) | -1.29%                          |
| redis (2 I/O threads)| 1.34%                           |
| redis (3 I/O threads)| 0.01%                           |

**Table 2.** We isolated OpenSSL keys in server applications with ERIM (ERIM-OpenSSL). We report the overhead of ERIM-OpenSSL with GARMR sandbox compared to the native execution. GARMR sandbox here refers to the sandbox described in Section 5.3.1.

We show the overhead for each successful experiment in Table 1. For the experiments, we removed ERIM’s sandbox (no sandbox) or replaced it with the sandbox described in Section 5.3.1 (GARMR sandbox). For ERIM-CPI with GARMR sandbox we report overhead of 2.23–6.20% with geometric mean of 4.10%, while for ERIM-SS with GARMR sandbox we report overhead of -0.03–3.32% with geometric mean of 1.43%. We also measured the overhead of standalone ERIM-CPI and ERIM-SS without the protection of any sandbox to show that even in the worst case (6.20%), most of the overhead (5.72%) can be attributed to the PKU-based memory isolation scheme and not the sandbox.

### 7.1.2 Isolating OpenSSL keys in server applications.

Similar to previous work [50], we isolated OpenSSL session keys in popular server applications with ERIM (ERIM-OpenSSL), to protect against server application vulnerabilities such as Heartbleed [18]. We configured `lighttpd`, nginx and redis, through their config files, to use ERIM-OpenSSL and only use ECDHE-RSA-AES128-GCM-SHA256 cipher and AES encryption for sessions. For the experiments, we replaced ERIM’s sandbox with the sandbox described in Section 5.3.1 (GARMR sandbox). Our results are shown in Table 2. We report an overhead of -1.29–1.34% with geometric mean of 0.47%.

| APP                  | XOM-Switch with GARMR Sandbox |
|----------------------|--------------------------------|
| nginx (1 worker)     | 0.04%                          |
| nginx (2 workers)    | -0.02%                         |
| nginx (3 workers)    | -0.09%                         |
| lighttpd (1 worker)  | 0.02%                          |
| lighttpd (2 workers) | 0.50%                          |
| lighttpd (3 workers) | 0.16%                          |
| redis (1 I/O thread) | 1.48%                          |
| redis (2 I/O threads)| 0.88%                          |
| redis (3 I/O threads)| 0.00%                          |

**Table 3.** We applied eXecute Only Memory (XOM) using XOM-Switch. We protected XOM-Switch with the GARMR sandbox. GARMR sandbox here refers to the sandbox described in Section 5.3.2. We report the overhead of XOM-Switch with this sandbox compared to the native execution.

We also measured the overhead of standalone ERIM-CPI and ERIM-SS without the protection of any sandbox to show that even in the worst case (6.20%), most of the overhead (5.72%) can be attributed to the PKU-based memory isolation scheme and not the sandbox.

### 7.1.3 Protecting Execute Only Memory.

We used XOM-Switch [39] to apply eXecute Only Memory (XOM). XOM-Switch is vulnerable to attackers that attempt to abuse PKU to disable XOM, unless it is combined with an additional mitigation such as CFI. We lifted this requirement by combining XOM-Switch with the PKU-based sandbox described in Section 5.3.2 (GARMR sandbox). Our results are depicted in Table 3. We report overhead of -0.09–1.48% with geometric mean of 0.33%.
| Attack                                    | GARMR Sandboxes |
|-------------------------------------------|-----------------|
| Inconsistencies of PT permissions [12]    | ✓               |
| Inconsistencies of PKU permissions [12]   | ✓               |
| Mapping with mutable backings [12]        | ✓               |
| Changing code by relocation [12]          | ✓               |
| Influencing intra-process behavior with seccomp [12] | ✓ |
| Modifying trusted mappings [12]           | ✓               |
| Race conditions in scanning [12]          | ✓               |
| Determination of trusted mappings [12]    | ✓               |
| Signal context attacks [12]               | X               |
| Vetted unsafe instruction relocation Section 4.3.1 | ✓ |
| Incomplete debug register update Section 4.3.2 | ✓ |

Table 5. Security analysis of the PKU-based sandboxes described in Section 5.3. We refer to these sandboxes as GARMR sandboxes. ✓ indicates that the sandbox stops the attack, while X indicates the opposite.

### 7.2 Security and Completeness

We analyzed the security of the constructed sandboxes on existing proof-of-concept attacks [12] and the two additional attacks that we discovered while building GARMR (see Section 4.3). For the former, we used the open source implementation of the exploits\(^8\), while for the latter we used our proof-of-concept exploits. We verified that the developed sandboxes stop all the attacks described in Sections 4.2 and 4.3, except for signal context attacks (see Section 4.2.7). This is not a fundamental limitation of our approach, but of the current prototypes of the framework and the constructed sandboxes. We discuss potential solutions to stop signal context attacks in Section 8. Our results are depicted in Table 5.

We compared the sandboxes described in Sections 5.3.1 and 5.3.2 (we refer to them as GARMR sandboxes) with the ones provided by ERIM and Hodor. The comparison is shown in Table 4. GARMR sandboxes are the first PKU-based sandboxes that can handle unsafe instructions without causing security or usability issues (see Section 4.1). GARMR sandboxes are also the first PKU-based sandboxes that prevent the attacks described in Sections 4.2 and 4.3, except for signal context attacks (see Section 4.2.7). The new attacks described in Section 4.3 specifically target Hodor’s instruction vetting mechanism, and are not applicable (N/A) to ERIM. Both GARMR sandboxes and ERIM’s sandbox are implemented in user space; nevertheless they require a small kernel patch to optimize performance. Last, all the sandboxes in this list incur low performance overhead.

### 8 Discussion

In this section, we discuss the limitations of our approach and alternative solutions.

**Signal context attacks.** A concurrent work with ours [7], proposes Endokernel, a subprocess virtualization scheme to deal with challenges in in-process isolation. Specifically for signals, the authors describe a signal virtualization mechanism that prevents attackers from tampering PKRU by abusing signals (see Section 4.2.7). However, two of the three provided implementations of Endokernel rely on additional mitigations, software diversity and CFI, which come with their own limitations regarding efficacy and performance.

Some of the techniques used in Endokernel could be applied to our framework to prevent signal context attacks. We believe that this would be feasible, since ptrace can intercept, delay, deny and redirect signals in ways similar to Endokernel [7].

**Restrictions on memory mappings.** GARMR restricts memory mappings to deal with attackers that target mappings with mutable backings as described in Section 5.3.1. This might lead to usability issues in applications such as older JIT engines, which used double-mapping as a way to bypass SELinux’s W\(^\ast\)X policy [17, 41]. Modern JIT engines no longer use double-mapping, however, since it can be detrimental to the application’s security [14].

**Multi-threaded applications.** Currently, we use instruction emulation to deal with applications that use multiple threads. However, this might add significant overhead in cases where several pages contain unsafe instructions. One suitable alternative could be to use process-wide events monitoring that will be available in future kernels [19]. Otherwise, we could implement a stop-the-world mechanism using signals to ensure that all the threads of a process are stopped.

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\(^8\)https://github.com/VolSec/pku-pitfalls
and then update their debug registers synchronously. This would require though to first provide a concrete solution for signal context attacks (see Section 4.2.7).

**Multiple levels of trust.** We only experimented with application partitioning schemes that use two levels of trust (trusted and untrusted). However, with our framework it would be possible to develop sandboxes for systems that use more than two memory domains, since the constructed sandboxes can reliably determine the current executing domain by inspecting PKRU as described in Section 5.1, and track the sensitive data of each domain.

9 Related Work

In-process isolation has been explored in depth, resulting in dozens of systems. In this section, we summarize works on in-process isolation that do not rely on PKU.

**OS abstractions.** Previous work introduced OS abstractions to enable multiple memory views and fast transition between them within a process address space [4, 11, 28, 36]. These approaches expose thread-like entities, control which resources they can access, and permit efficient transitions between them. However, all these techniques are not directly applicable to legacy code without code modifications.

**Virtualization-based techniques.** Hodor [27] and SeCage [37] leverage virtualization extensions [13] (VT-x), to provide different memory views for trusted and untrusted code. SeCage is vulnerable to an in-process adversary unless it is used in conjunction with CFI, while Hodor’s VT-x based implementation is less efficient than Hodor’s PKU-based counterpart [27]. Intel and AMD CPUs provide Supervisor-mode Access Prevention (SMAP) hardware feature to disable kernel accesses to user space memory. Seimei uses SMAP and VT-x to provide low cost and secure in-process isolation [56]. xMP [45] extended Xen hypervisor’s a1tp2m subsystem [34, 44] and the Linux memory management system to isolate sensitive user space and kernel data in disjoint xMP memory domains.

**Hardware extensions.** Researchers also proposed hardware extensions to provide efficient fine-grained component isolation. CHERI [58] and CodomMs [54] extended the RISC and x86 ISAs respectively, with capabilities. Donky [46], on the other hand, augmented the x86 and RISC-V ISAs to provide secure memory protection domains similar to PKU. MicroStache [40] and IMIX [22] extended the x86 ISA with instructions to access safe regions. ARM memory domains [3] are similar to PKU domains, but they are only available on 32-bit chips and domain permissions can only be modified with privileged instructions. This paper focuses on solutions that can be built on commodity x86 CPUs.

**SFI.** Software fault isolation (SFI) restricts parts of an application code from accessing memory outside of designated bounds [10, 15, 16, 20, 21, 38, 48, 55, 59, 60]. SFI techniques employ complex static and dynamic analysis and instrumentation that introduce non-negligible overhead. In addition, many of the proposed techniques rely on an additional mitigation such as CFI, to prevent in-process attackers from bypassing bounds checks.

**Compartmentalization.** Partitioning an application into compartments and defining which resources they can access is an open problem and it is orthogonal to this paper. Previous work focuses on identifying suitable isolation boundaries in applications using automatic and semiautomatic (e.g., annotations) techniques [6, 24, 26, 52, 53]. However, completely automating compartmentalization of existing software is still challenging.

10 Conclusion

Recent research has explicitly highlighted the extreme care that should be taken when implementing PKU-based sandboxing, mentioning a large number of edge cases and a difference in perspective between the OS and the security community on PKU as contributing factors. In this paper, we analyzed the various challenges of PKU-based sandboxing. We also introduced two new proof-of-concept attacks that target Hodor and that manage to bypass its sandbox.

We then presented GarMr, a new PKU-based sandboxing framework that facilitates development of PKU-based sandboxes. We applied our framework to build sandboxes for two state-of-the-art PKU-based memory isolation systems: ERIM and XOM-Switch. We evaluated the security and performance of the constructed sandboxes using proof-of-concept exploits and high-performance server applications respectively. Our extensive evaluation shows that GARMR overcomes limitations of existing work, enabling practical, efficient and secure PKU-based sandboxing.

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