I. INTRODUCTION

As many effective measures for protecting call stacks get deployed (such as canaries [1], reordering local variables [2], and Safe SEH [3]), heap vulnerabilities gain growing attention of attackers. Heap vulnerabilities can be exploited by attackers to launch vicious attacks. The recent Heartbleed [4] and WannaCry [5] attacks demonstrate the dangers. For instance, the WannaCry ransomware uses the EternalBlue exploit, which makes use of a heap buffer overwrite vulnerability to hijack the control flow of the victim program [5].

There are a variety of heap vulnerability types. The following types are among the most commonly exploited types.\(^1\) (1) **Buffer overflow**: it includes both *overwrite* and *overread*. By overwriting a buffer, the attack can manipulate data adjacent to that buffer and launch various control-data or non-control-data attacks, while exploitation of overread can steal sensitive information in memory, such as address space layout and private keys. (2) **Use after free**: it refers to accessing memory after it has been freed. If the memory space being reused is under the control of attackers, use-after-free bugs can be exploited to launch various attacks, such as control flow hijacking. (3) **Uninitialized read**: exploitation of such vulnerabilities can leak sensitive information.

\(^1\) *Double free* was frequently exploited; but many popular allocators, such as the default allocator in glibc [6], have built-in double free detection now.

Many approaches have been proposed to tackle heap vulnerabilities. Some systems try to discover zero-day heap vulnerabilities before software release [7]–[9]. Yet, it is very unlikely to find all of them. A large body of research focuses on detecting, preventing or mitigating heap attacks (and other memory-based attacks) [10]–[26]. They usually incur a large overhead and/or can only handle a specific type of heap vulnerabilities. For example, MemorySanitizer [22] is a dynamic tool that detects *uninitialized read*; however, it incurs 2.5x of slowdown and 2x of memory overhead. AddressSanitizer [10], which detects *overflows* and *use after free* online, is deemed fast, but still incurs 73% slowdown and 3.4x memory overhead. As another example, HeapTherapy [21] proposes an efficient *heap buffer overflow* detection and response system; however, it does not provide methods for detecting and handling *uninitialized read* and *use after free*.

When examining the spectrum of heap security measures, we notice that handling heap vulnerabilities through *patching* has been much less studied. Patching, however, has been an indispensable step for handling vulnerabilities in practice. Over decades, conventional patch generation and deployment have suffered serious limitations. First, the patch generation is a lengthy procedure. Even for security sensitive bugs, it takes those big vendors 153 days on average from vulnerability report to patch availability [27]. A study finds that only 65% of vulnerabilities in software running on a typical Windows host have patches available at vulnerability disclosure [28]. This provides opportunities for attackers to exploit the unpatched vulnerabilities on a large scale [29]. For resource-constrained small software companies, it takes even longer time. Plus, it is unlikely to generate patches for legacy software, whose support from the vendor has ended.

Second, given a vulnerability, its fresh patches may have not been thoroughly tested, and thus tend to introduce stability issues and even logic errors. Although waiting for mature patches can reduce the risk, it makes the exploitation window longer. This has been a dilemma in patch deployment [30].

We propose a heap patching system that does not have the limitations above. Our insight is that, by changing the configuration of heap memory allocation, *all* the aforementioned heap vulnerabilities can be addressed without altering the program code and, hence, no new bugs are introduced. Based on the configuration information, the allocator can accordingly enhance its handling (i.e., allocation, initialization and deallocation) of buffers that are vulnerable to attacks,
called vulnerable buffers, and apply security enhancement only
to them (rather than all heap buffers) to minimize the overhead.
We thus propose to generate Heap Patches as Configuration
and call our system HPAC.
HPAC consists of a heavyweight offline patch generation
phase and a lightweight online defense generation phase. In
the offline patch generation phase, we use shadow memory to
scrutinize attacks and achieve byte precision level. We group
buffers according to their allocation-time calling contexts. Buffers that share the same allocation-time calling context as
the buffer exploited by the attack are regarded as vulnerable
buffers. The allocation-time calling context of vulnerable buffers along with other information is collected to generate
patches, i.e., the configuration information. Next, in the online
defense generation phase, the configuration information is
loaded and the stored calling context information guides the
allocator to recognize vulnerable buffers. It properly combines
detailed offline analysis and highly efficient online defenses.
However, if call stack walking (as used by gdb) is used for
obtaining calling contexts, it can incur significant slowdown,
especially for allocation-intensive programs [31]–[33]. We
thus use calling context encoding, which continuously rep-
resents the current calling context in one or a few integers [31].
By reading the integer(s), the encoded calling context, called
Calling Context ID (CCID), can be obtained. By comparing
the CCID for the current buffer allocation with the CCIDs
stored in the configuration information, the online system
can swiftly determine whether the new buffer is vulnerable.
Moreover, we propose targeted calling context encoding,
which is a suite of algorithms that can optimize many famous
calling encoding methods, such as PCC [31], PCCE [32], and
DeltaPath [33]. Since calling context encoding is an important
technique with many applications, the optimization algorithms
constitute a separate contribution.

Installing a heap patch does not change the program code.
Specifically, a heap patch is in the form of a (key, value)
tuple, where the key is the allocation-time CCID of the
vulnerable buffer and the value indicates the vulnerability type
and the parameter(s) for applying the online defense. The
patches are read into a hash table upon program initialization.
It thus takes only O(1) time to determine whether a new buffer
is vulnerable. The online defense is enforced by intercepting
heap buffer allocation and deallocation. Both the hash table
initialization and the buffer allocation/deallocation interception
are transparent to the underlying heap allocator, and imple-
mented in a shared library.2 We thus do not need to change the
underlying heap allocator or depend on a specific allocator.

None of the techniques used in HPAC, except for targeted
calling context encoding, is new. However, static analysis, code
instrumentation, offline attack analysis, and online defense
generation are creatively combined to build a new counter-
measure against heap attacks. A comprehensive evaluation is
performed, showing that HPAC is effective and efficient. We
make the following contributions.

2In Linux, we can load it using LD_PRELOAD.

• We properly combine heavyweight offline attack analysis
and lightweight online defense generation to build a new
heap defense system that simultaneously demonstrates
the following good properties: (1) patch generation with-
out manual efforts, (2) code-less patching, (3) versatile
handling of heap buffer overwrite, overread, use after
free, and uninitialized read, (4) imposing a very small
overhead, and (5) no dependency on specific allocators.
• We propose targeted calling context encoding, a suite of
algorithms that can optimize calling context encoding,
and demonstrate its application to our system.

II. RELATED WORKS

Given the large body of research on heap memory safety, we
do not intend to make an exhaustive list of work on the prob-
lem. Instead, we compare HPAC with other automatic patch
generation techniques, and then examine critical techniques
used in our system.

A. Automatic Patch/Defense Generation

With attack inputs in hand, generating patches/defenses au-
tomatically has been a highly desired goal. We divide previous
researches towards this goal into the following categories.

Bytes pattern based signature generation. Given a large
number of attack inputs, many systems (such as Honey-
comb [34], Autograph [35], and Polygraph [36]) generate sig-
natures by extracting common bytes patterns from the inputs.
However, such methods usually need many attack samples in
order to correctly mine patterns, and cannot work when only
one or very few attack inputs are available. False positives may
be raised when benign inputs happen to match the signatures.
Plus, attackers can mutate the inputs to bypass the detection.
In addition, these systems usually have deployment difficulty
in handling compressed or encrypted inputs.

Semantics based signature generation. Tools like COV-
ERS [37], Hamsa [38], TaintCheck [39] and the work by
Xu et al. [40] propose methods to generate semantics-based
signatures; e.g., spotting the target system call ID used upon
control flow hijacking and filtering out inputs that contain that
ID. They are very effective in handling certain control flow
hijacking attacks, but it is unknown how they can be applied
to addressing overread and uninitialized read. They also have
deployment difficulty in handling compressed and encrypted
attack inputs and may incur false positives.

Tracking faulty instructions. By replaying the attacks, some
systems try to pinpoint faulty instructions that are exploited
by the attacks and try to generate patches to fix them; such
systems include VSEF [41], Vigilant [42], PASAN [43] and
AutoPag [44]. A frequently employed insight is that a tainted
input, e.g., due to overwrite, should not be used to calculate
the indirect jump address. It is unknown how such systems
can handle attacks beyond control flow hijacking, e.g., buffer
overread attacks. Plus, the deployment of the patches requires
code update, just like conventional code patching.

Trial and error for patch generation. Some systems propose
genetic programming based program generation [45], template
recently, calling context is used to generate context sensitive defenses [21], [40], [41], [53], [54]. In particular, Exterminator [53] also proposes to generate context-sensitive heap patches. However, our system HPAC differs from Exterminator in multiple aspects. (1) Exterminator performs online probabilistic attack detection (e.g., when an overflow occurs, it may or may not detect it), while HPAC performs offline deterministic attack analysis and patch generation. How to apply patches generated by heavyweight offline analysis to lightweight online defense generation is not trivial and solved by our work. (2) Exterminator does not handle overread or uninitialized-read, while HPAC handles all the frequently exploited heap vulnerability types including overwrite, overread, use after free, and uninitialized read. (3) Exterminator relies on a custom heap allocator that incurs large overheads, while HPAC does not; the defense of HPAC is transparent to the underlying allocator. (4) Exterminator uses the expensive stack walking to retrieve calling contexts, while targeted calling context encoding is proposed and applied in HPAC. But the two works share the insight in calling context-sensitive heap patches, which we do not claim as our contribution.

B. Calling Context Encoding

Background. A calling context is the sequence of active function calls on the call stack. It carries critical information about dynamic program behavior. It thus has been widely used in debugging, testing, anomaly detection, event logging, performance optimization, and profiling [33]. For example, logging sensitive system calls is a practice in many systems. Recording the calling context of the system call provides important information about the sequence of program components that gets involved and leads to the call.

Obtaining calling contexts through stack walking is straightforward but very expensive [31]. A few encoding techniques, which represent a calling context using one or very few integers, have been proposed to continuously track calling contexts with a low overhead. The probabilistic calling context (PCC) technique [31] computes a probabilistically unique integer ID, essentially a hash value, for each calling context, but does not support decoding. Precise calling context encoding (PCCE) [32] stems from path profiling [50] and supports decoding. Another example is DeltaPath [33], which improves PCCE by supporting virtual function calls and large-sized programs. A relevant but different problem is path encoding [50], which represents program execution paths (within a control flow graph) into integers.

Similar to targeted calling context encoding, another work [51] also aims to minimize the overhead due to the encoding, but uses a very different idea. It performs offline-profiling runs to establish the mapping between stack offsets and calling contexts. It fails if the calling context of interest does not appear in the profiling runs. Its reported decoding failure rate is as high as 27%. Finally, it does work if variable-size local arrays (allowed in C/C++) are used.

C. Calling Context-Sensitive Defenses

Calling context was applied to areas beyond debugging decades ago. As an example, a region-based heap allocator tags heap objects with allocation-time calling context [52].
heap vulnerabilities, patches only need to be written into a configuration file $C$ to take effect.

B. System Architecture

As shown in Figure 1, the system consists of the following components: (1) A Program Instrumentation Tool: it builds the calling context encoding capability into the program (Section IV). Note that the program instrumentation is an one-time effort. Because of the simplicity of the instrumentation, its correctness can be verified automatically. The instrumented program is then used for both offline patch generation and the online system. (2) An Offline Patch Generator: it automatically generates the patch by replaying the attack (Section V). (3) An Online Defense Generator: it is a dynamically linked library that (a) loads the patches from the configuration file $C$ at program start, and (b) intercepts buffer allocation operations for recognizing vulnerable buffers and generate security measures online (Section VI).

C. Calling-Context Sensitive Patches

Given the attack input that exploits a heap vulnerability $V$, in order to generate a patch $P$ based on attack analysis, it is critical to extract some invariant among attack instances. Such invariant then can be used to design protection against future attacks that also exploit $V$.

Our observation is that attacks that exploit $V$ usually share some attack-time calling context (e.g., the sequence of active function calls that lead to a buffer overflow due to a `memcpy` call). If we trace the program execution backward, these vulnerable buffers probably share the allocation-time calling context, which we call a vulnerable calling context and can be used as an invariant to generate the patch $P$. Whenever a heap buffer is allocated, the current calling context is retrieved and compared with vulnerable calling contexts to determine whether the buffer being allocated is vulnerable.

IV. TARGETED CALLING CONTEXT ENCODING

Simple call stack walking for retrieving calling contexts would incur a large overhead, especially for programs with intensive heap allocations [31]. There exist several efficient calling context encoding techniques that are famous, such as [31]–[33]. We propose targeted calling context encoding, which is a suite of algorithms that can be used to optimize these encoding techniques. The insight is that when the target functions, whose calling contexts are of interest, are known, many irrelevant call sites do not need to be instrumented and thus the overhead can be significantly reduced.

The input of our algorithms is the call graph of the program, and the output is a pruned sub-call-graph. As each of the three famous encoding techniques [31]–[33] can take a graph as input, they all should work well when the sub-call-graph if fed to these systems. To make the discussion concrete (and based on our choice of the encoding technique for heap patching), we use Probabilistic Calling Context (PCC) [31] to demonstrate the application of the proposed optimizations.

According to PCC, at the prologue of each function, the current calling context ID (CCID), which is stored in a thread-local integer variable $V$, is read into a local variable $t$; right before each call site, $V$ is updated as $V = 3 \ast t + c$, where $c$ is a random constant unique for each call site.\footnote{The encoding in PCCE [32] and DeltaPath [33] basically adopts $V = t+c$, where $c$ is calculated according to the encoding algorithms.} This way, $V$ continuously stores the current CCID. Thus, the current CCID can be obtained conveniently by reading $V$. With PCC, however, it may happen that multiple calling contexts obtain the same encoding due to hash collisions. It is shown practically and theoretically that the chance of hash collision is very low [31]. It is worth noting that a hash collision in our system means that a non-vulnerable buffer is recognized as a vulnerable buffer and gets enhanced. Any of our enhancements do not change the program logic, so a hash collision can cause unnecessary overhead, but it does not affect the correctness of our system.
We call the original encoding algorithm that take all the call sites into consideration as Full-Call-Site (FCS) instrumentation. The three famous encoding algorithms, PCC [31], PCCE [32] and DeltaPath [33] all enforce FCS. Figure 2(a) shows that all the call sites in those red nodes are instrumented, and $T_1$ and $T_2$ are the target functions. The less call sites are instrumented, the smaller overhead is expected.

A. Targeted-Call-Site (TCS) Optimization

FCS blindly instruments all the call sites in a program. In practice, very often users are only interested in the calling contexts that end at one of a specific set of target functions, such as security-sensitive system calls and critical transaction calls. In our case, we are only interested in calling contexts when the allocation APIs (such as malloc, calloc, memalign, aligned_alloc) are invoked. It is unnecessary to instrument functions that may never appear in the call stacks when these target functions are invoked.

We thus propose the first optimization, Targeted-Call-Site (TCS), where only the call sites that may appear in the calling contexts of target functions are instrumented. To conduct the TCS optimization, reachability analysis on the call graph of the program is performed. Given a call graph $G = (V, E)$, where $V$ is the set of nodes representing functions of the program and $E$ the set of function calls, and a set of functions $F$, we perform reachability analysis to find edges that can reach any of the functions in $F$, and only call sites corresponding to these edges are instrumented.

Figure 2(b) shows the instrumentation result of TCS. As the edges $DH$ and $HI$ cannot reach any of the target functions $T_1$ and $T_2$, they are pruned from the instrumentation, reducing the set of call sites that need to be instrumented.

B. Slim Optimization

On the basis of TCS, there is still potential to further prune the set of call sites to be instrumented. In a call graph, a node can be classified as either a branching or non-branching one: a branching node is one that has multiple outgoing edges that can reach (one of) the target functions. Our insight is that the purpose of call site instrumentation is to make sure different calling contexts can obtain different encoding values; given a non-branching node, whether or how its contained call sites are instrumented does not affect the distinguishability of the encoding results. Thus, we propose to avoid instrumenting the call sites in those non-branching nodes.

For example, as shown in Figure 2(c), according to the Slim optimization, all call sites in the non-branching nodes, $B$ and $E$, are excluded from the instrumentation set.

C. Incremental Optimization

The two optimization algorithms treat all target functions as a whole. Our another insight is that when the call to a target function is intercepted for analysis or logging purpose, the analyzer or logger usually knows the target function. In our case, when malloc and memalign are intercepted, different interception functions will be invoked.

Therefore, we can use the pair of $\{\text{Target\_fun, CCID}\}$ (rather than CCID alone) to distinguish different calling contexts. Based on this insight, we propose another optimization algorithm that can further reduce the number of instrumented call sites. A node is an true branching node if it has two or more outgoing edges that reach the same target function. That is, if a node has multiple outgoing edges, each of which reaches a different target function, it is called a false branching node. The idea of the Incremental encoding is to avoid instrumentation the call sites in a false branching node.

In Figure 2, node $A$ is a true branching node, as its two outgoing edges can reach the same node $T_1$ (and $T_2$ as well). So is node $C$, as its two outgoing edges can reach $T_1$. Thus, only the call sites that correspond to $AB$, $AC$, $CE$, $CF$ need to be instrumented. Take the calling contexts of $T_2$ as an example, the instrumentation at $AB$ and $AC$ is sufficient to distinguish the two calling contexts that reach $T_2$.

Algorithm 1 Incremental Optimization.

**Input**: A call graph $CG = (N, E)$, and the set of target functions $T \subseteq N$.

**Output**: The functions in $N$ to be instrumented.

1: function FILTER($T, CG = (N, E)$):
2:   $\text{InstrumentationSet} \leftarrow \{\}$
3:   for each $t \in T$ do
4:     $\text{VisitedNodes} \leftarrow \{\}$
5:     $\text{Queue}$.push($t$)
6:     while $\text{Queue}$.size() > 0 do
7:       $n \leftarrow \text{Queue}.pop()$
8:       for each $e = (m, n)$ of the incoming edges of $n$ do
9:         if $m \notin \text{VisitedNodes}$ then
10:            $\text{Queue}.push(m)$
11:     for each $e = (n, m)$ of outgoing edges of $n$ do
12:       if $m \notin \text{VisitedNodes}$ then
13:         count $\leftarrow$ count + 1
14:     if count $>$ 1 then
15:       $\text{InstrumentationSet}.push(n)$
16:   return $\text{InstrumentationSet}$

Algorithm 1 shows the algorithm for incremental optimization. Line 3 illustrates the idea of processing each target function incrementally. For each target function, Lines 4–17 are to find true branching nodes relative to it. Specifically, Lines 4–10 are a backward breadth-first search; as it omits nodes already visited (Line 9), it can correctly handle back edges. Then Lines 11–17 are to find true branching nodes.

In short, the three encoding optimization algorithms are based on different insights and ideas, and each improves on the previous one in terms of reducing the set of call sites to be instrumented.

V. Offline Attack Analysis and Patch Generation

The Offline Patch Generator component runs the vulnerable program using the attack input and generates the patch as part of the dynamic analysis report. It is built on dynamic binary instrumentation and shadow memory of Valgrind [55]. As shown in Figure 3, for every bit of the program memory,
typedef struct {
    uint32_t i;
    uint8_t c;
} A;

A y, *p = (A *) malloc( sizeof(A) );
p->i = 0; p->c = 'f';
y = *p;

Fig. 4. Legal uninitialized read due to padding.

a Validity bit (V-bit) is maintained to indicate whether the accompanying bit has a legitimate value; instructions are inserted for the propagation of V-bits when data copy occurs (e.g., when a word is read from memory to a register); for every byte of the memory location, an Accessibility bit (A-bit) is maintained to indicate whether the memory location can be accessed.

When a heap buffer is malloc-ed, the returned memory is marked as accessible but invalid. Each buffer is surrounded by a pair of red zones (16 bytes each), which are marked as inaccessible. When a heap buffer is free-ed, its memory is set as inaccessible. In addition, whenever a heap buffer is allocated, the current calling context ID (CCID) is recorded and associated with the buffer.

1) Detecting overflows: A buffer overflow will access the inaccessible red zone appended to the buffer and get detected.

2) Detecting use after free: A free-ed buffer is set as inaccessible and then added to a FIFO queue of freed blocks. Thus, the memory is not immediately made available for reuse. Any attempts to access any of the blocks in the queue can be detected. The maximum total size of the buffers in the queue is set as 2GB by default, which is large enough for the exploits we investigated, and can be customized. In Section IX, we discuss how to handle it if the quota is insufficient.

3) Detecting uninitialized read: To detect uninitialized read, an attempt is to report any access to uninitialized data, but this will lead to many false positives. For instance, given the code snippet in Figure 4, most of the compilers will round the size of A to 8 bytes; so only 5 bytes of the heap buffer is initialized (and the V-bits for the remaining 3 bytes are zero), while the compiler typically generates code to copy all 8 bytes for y = *p, which would cause false positives due to accessing the 3 bytes whose V-bits are zero.

To avoid false positives due to padding, we check the V-bit of a value only when it is used to decide the control flow (e.g., jnz), used as a memory address, or used in a system call (as the kernel behavior is not tracked). As every bit of the program has a V-bit, bit-precision detection of uninitialized read is achieved. Moreover, origin tracking is used to track the use of invalid data back to the uninitialized data (such as a heap buffer) when a warning is raised, which allows us to retrieve the allocation-time CCID associated with the vulnerable buffer. When an attack is detected, the patch is generated in the form of ⟨FUN, CCID, T⟩, where FUN is the function used to request the heap buffer (such as malloc, memalign), CCID is an integer representing the allocation-time calling context ID of the vulnerable buffer, and T is a three-bit integer representing the vulnerability type (the three bits are used to indicate OVERFLOW, USE-AFTER-FREE, UNINITIALIZED-READ, respectively). Example patches are shown in the upper graph in Figure 5.

How to handle realloc: If the new size is smaller than the original size, the cut-off region is marked as inaccessible. If the new size is larger, the added region is set as accessible but invalid. The allocation-time CCID associated with the buffer is also updated with the value of the realloc invocation.

How to handle multiple vulnerabilities: An attack input may exploit multiple vulnerabilities. For example, the Heartbleed attack exploits both uninitialized read and overread. In order to handle the case that an attack exploits multiple vulnerabilities, we resume the program execution upon warnings. Plus, once the V bits for a value have been checked, they are then set to valid; this avoids a large number of chained warnings. Finally, a script is used to process the many warnings according to the origin (i.e., the address of the vulnerable buffer) of those warnings and generate patches correctly.

VI. CODE-LESS PATCHING AND ONLINE DEFENSES

When the patched program is started, as shown in Figure 5, the Online Defense Generator library has an initialization function that reads patches from the configuration file and stores them into a hash table, where the key of each entry is ⟨ALLOCATION_FUNCTION, CCID⟩ and the value is the vulnerability type(s) and parameters, if any, for applying the security measures. Note once the hash table is initialized, its memory pages are set as read only.
The Online Defense Generator library intercepts all heap memory allocation operations. Whenever a heap buffer is allocated, the name of the allocation function (hardcoded into the interposing function) along with the current CCID is used to search in the patch hash table, which takes only O(1) time. If there is no match, the buffer does not need to be enhanced; otherwise, the buffer is enhanced based on the associated vulnerability type(s) and parameters.

While the security measures themselves are straightforward, several considerations make the design challenging. (1) In some cases, the same buffer may be vulnerable to multiple attacks, such as uninitialized read and overflow. (2) In addition to handling malloc and free, the system needs to support a family of other allocation functions, such as realloc and memalign (aligned allocation). These challenges are well resolved by our system.

One complexity is that we maintain heap metadata ourselves, such as the buffer size (to support realloc correctly), vulnerability type(s), the buffer alignment information, and the location of the guard page, so that our system can work without having to change the underlying allocator or rely on its internals.

(1) Handling overflows: If the buffer is vulnerable to overflows, a guard page is appended to it to prevent such attacks. While the guard page can effectively prevent overflows, they are known to be prohibitively expensive when being applied to every buffer. In our system, however, the guard page is precisely applied to vulnerable buffers, and the resulting overhead is dramatically reduced.

As shown in Figure 6, Structure 2 is used for non-aligned buffers, while Structure 4 is used for aligned buffers (allocated using memalign, etc.). When a heap allocation request is intercepted, the requested size is increased to accommodate the word for metadata and the guard page (as well as necessary padding following the user buffer to ensure the guard page is page aligned). The address of the user buffer is returned to service the user program.

The metadata word contains rich information and is worth detailed interpretation. (1) In all structures, the least significant four bits is called the buffer type field, where three bits represent the vulnerability type (one bit is used to indicate each of the three vulnerability types, i.e., Overflow, Use after Free, and Uninitialized Read) and one bit indicates whether the buffer is aligned. (2) 36 bits are used to indicate the location of the guard page. Currently, 64-bit operations systems only use a 48-bit virtual address space; plus, a guard page is $4KB = 2^{12}B$ aligned. Thus, $48 - 12 = 36$ bits are sufficient. A guard page is set as inaccessible using mprotect. The user buffer size information is stored as the first word of the guard page, and it is needed for supporting realloc. (3) If the buffer is aligned (Structure 3 and Structure 4), there is a padding field whose size depends on the alignment size. The alignment size information is needed to determine the buffer address given the address of the User Buffer upon a free call. As the alignment size is always a power of two (i.e., $2^n$), we only need 6 bits to store the value of $n \in [0, 64]$, which then can be used to calculate the alignment size.

(2) Handling use after free: If an allocation is not aligned, the buffer takes Structure 1; otherwise, Structure 3. The metadata word uses 48 bits to store the user buffer size. When a buffer vulnerable to use after free is to be free-ed, it is put into an FIFO queue of freed blocks to defer the reuse. In our system, only buffers vulnerable to use-after-free are put into the queue, such that given the same quota the time a freed buffer stays in the queue is much lengthened, which hence significantly increases the difficulty of exploitation of a use-after-free vulnerability for it increases the uncertainty entropy a freed buffer is reused by attackers.

(3) Handling uninitialized read: Similar to the above, if the allocation is not aligned, the buffer takes Structure 1; otherwise, Structure 3. The user buffer region is initialized with zeros before it is returned to the user program.

Table I summarizes how different buffer structures are used for handling different cases, including when multiple vulnerabilities affect the same buffer. If there is a threat of overflow, Structure 2 or Structure 4 is used to accommodate the guard page depending on whether the allocation call is memalign. Whenever there is use after free, upon being freed the buffer is put into the freed-blocks queue to defer the reuse of these buffers.

How the Online System Handles free() calls: A particular advantage of our system is that it supports the deployment of heap patches without modifying the underlying allocator. It works solely by intercepting the memory allocation calls. On
TABLE I
A SUMMARY OF THE USE OF BUFFER STRUCTURES.

| Vulnerability type       | Not aligned | Aligned |
|--------------------------|-------------|---------|
| Not Vulnerable           | Structure 1 | Structure 3 |
| Overflow                 | Structure 2 | Structure 4 |
| Use-after-free           | Structure 1 | Structure 3 |
| Uninitialized read       | Structure 1 | Structure 3 |
| Overflow & Use-after-free| Structure 2 | Structure 4 |
| Uninitialized read       | Structure 2 | Structure 4 |
| Use-after-free & Uninitialized read | Structure 1 | Structure 3 |
| Overflow & Use-after-free & Uninitialized read | Structure 2 | Structure 4 |

the other, it complicates the handling of freeing buffers.

As shown in Figure 7, when free(p) is invoked by the user program, the Online Defense Generator intercepts the call and handles it as follows. (1) If the Overflow bit in the metadata word is set, the location information of the guard page is retrieved and the guard page is set as accessible using mprotect. (2) Based on the user buffer address p, the initial address of the buffer pi is calculated. Specifically, if the buffer was not allocated using memalign, pi = p - sizeof(void*); otherwise, the alignment size A is retrieved and pi = p - A. (3) If the Use-after-Free bit is set, the block is put into the queue of the freed blocks; otherwise, the buffer is released using the original free API of the underlying allocator and the buffer address pi is passed to the call.

VII. OTHER IMPLEMENTATION DETAILS

Implementation of the Program Instrumentation Tool. Currently, we add a pass into LLVM, which performs the call graph analysis to determine the set of call sites to be instrumented and then instruments them. This implementation assumes the program source code is available. Another viable implementation path is based on binary code instrumentation, e.g., via Dyninst, which can insert code into the program dynamically. It is worth mentioning that Dyninst does not require the source code to perform instrumentation. Thus, software users (not only software companies) can also instrument their software (only once) and generate patches themselves.

Implementation of the Offline Patch Generator. This component is built on the basis of Valgrind [55]. We reuse its shadow memory functionality and modify the tool to support the needed handling of allocation and deallocation. Significant effort has been saved by making use of Valgrind, which in the meanwhile is a mature dynamic analysis tool. The implementation over Valgrind also benefits us to analyze various complex real-world programs successfully.

Implementation of the Online Defense Generator. It is implemented as a shared library, which reads the patches in the configuration file to the hash table and intercepts all the allocation function calls to enhance vulnerable buffers according to the patches. The hash table memory pages are set as read-only once the initialization is done. The library is loaded using LD_PRELOAD in our prototype. It does not change the underlying heap allocator or rely on its internals.

VIII. EVALUATION

We have evaluated HPAC in terms of both effectiveness and efficiency. We not only evaluate it on the SPEC CPU2006 benchmarks and many vulnerable programs, but also run the system with real-world service programs. The efficiency improvement of the calling context encoding optimization algorithms is also measured. Our experiments use a machine with a 2.8GHZ CPU, 16G RAM running 16.04 Ubuntu and Linux Kernel 4.10.

A. Effectiveness

To evaluate the effectiveness of our system HPAC, we run it on a series of programs, as shown in Table II, which contain a variety of heap vulnerabilities. In the effectiveness experiments, we aim to evaluate (1) whether the Offline Patch Generator can correctly determine the vulnerability type and generate patches; and (2) whether the generated patches can effectively prevent attacks from exploiting those heap vulnerabilities.

So far, our Dyninst-based solution only supports single threaded programs.
vulnerabilities. It is worth mentioning that HPAC, as a single system, can handle the variety of heap vulnerabilities; plus, different from conventional code patches, installing patches do not modify any line of the program code and hence do not introduce stability or logic errors (but we do require one-time program instrumentation). Below we describe details for some of the experiments we performed.

Heartbleed Attacks. Heartbleed was a notorious vulnerability of OpenSSL and affected a large number of services [58]. By sending an ill formed heartbeat request and receiving the response, the attacker can steal data from the vulnerable services, such as private keys and user account information. While Heartbleed is widely known as a heap buffer over-read vulnerability, actually the attacker can exploit two different vulnerabilities: over-read and uninitialized read. Specifically, the vulnerable heap buffer has 34KB, while the size \( l \) of the data stealing from the buffer can be up to 64KB. If \( l < 34K \), the attack is just an uninitialized read that leaks data previously stored in the buffer; otherwise, it is a mix of uninitialized read and over-read [59].

A service was created using the OpenSSL utility `s_server`.\(^6\) We then collected different attack inputs from Internet, and used one of them to generate the patch. Our Offline Patch Generator correctly identified it as a mix of uninitialized read and overflow and output the patch. The patch was then automatically written into the configuration file of the Online Defense Generator, which was able to precisely recognize and enhance the vulnerable buffers. We then tried different attack inputs, and no data was leaked except for the zeros filled in the buffers.

bc-1.06. `bc`, for basic calculator, is an arbitrary-precision calculator language with syntax similar to the C programming language. Some versions of its implementation contain a heap buffer overflow vulnerability. We obtained a buggy version of this program from BugBench, a C/C++ bug benchmark suite [56], and collected a malicious input that overflows buffers and corrupts the adjacent data. By feeding the input into our Offline Patch Generator that ran the buggy program, an overflow patch was generated. With the patch deployed, our system successfully stopped the attack before it corrupted any data.

GhostXPS 9.21. GhostXPS is an implementation of the Microsoft XPS document format built on top of Ghostscript, which is an interpreter/renderer for PostScript and normalizing PDF files. It is the leading independent interpreter software with the most comprehensive set of page description languages on the market today. Some versions of GhostXPS contain an uninitialized read vulnerability that can be exploited using a crafted document. We collected a buggy version of GhostXPS from their git repository and the malicious document input. In the offline patch generation phase, the uninitialized read attack was detected and a patch was generated. During the online heap protection phase, the attack was not able to steal any data, except for zeros, from memory.

optipng-0.6.4. OptiPNG is a PNG image optimizer that compresses image files to a smaller size without losing any information. Specific versions of this optimizer allow the attacker to exploit a use-after-free vulnerability and execute arbitrary code via crafted PNG files. We collected a vulnerable version (optipng-0.6.4) and a malicious PNG image. The Offline Patch Generator correctly identified the attack and generated a patch. The Online Defense Generator made use of the patch to recognize the vulnerable buffers and defeated the use-after-free attacks by deferring the deallocation of vulnerable buffers.

tiff-4.0.8. TIFF provides support for "Tag Image File Format", commonly used for sorting image data. In LibTIFF 4.0.8, there is a heap buffer overflow in the `t2p_write_pdf` function in tools/tiff-2pdf.c. We were able to generate the patch, which could successfully prevent the overflow.

SAMATE Dataset. We evaluated our system on the SAMATE Dataset, which is maintained by NIST [57] and contains 23 programs with heap buffer overflow, uninitialized read, or use after free vulnerabilities. Our system successfully generated patches for all of them and prevented the vulnerabilities from being exploited.

B. Efficiency

We compared the overhead incurred by the different calling context encoding algorithms, and measured the overall speed overhead and memory overhead incurred by our system. We used our LLVM-based implementation to measure the efficiency of different calling context encoding algorithms.

1) Overhead Comparison of Different Calling Context Encoding Algorithms: To measure the execution time overhead imposed by different calling context encoding algorithms, we applied them to the programs in the SPEC CPU2006 Integer benchmarks, and measured the execution time when different encoding techniques were applied, normalized using the execution time when no encoding is applied. Compared to FCS (Full Call-Site Instrumentation) proposed in [31], which incurred 2.4% of slowdown for C/C++ programs, the other three encoding algorithms proposed by us, that is, TCS (Targeted Call-Site Instrumentation), Slim, and Incremental,

\(^6\)In order to support the interposition of the allocation operations, we compiled OpenSSL using `OPENSSL_NO_BUF_FREELIST` compilation flag to disable the use of freelists.
incurred 0.6%, 0.5%, and 0.4% of slowdown, receptively. While the saved execution time itself is small, it gains up to 6x of speed up. We believe the proposed encoding algorithms can have many applications far beyond memory protection; plus, when they are applied to Java programs, where FCS may incur more than 35% of overhead [33], the speed up due to our algorithms could make a significant difference.

As the encoding works by inserting instructions into the programs, we also measured the program size increase. The results are shown in Table III. While FCS increased the binary size by an average of 12% when compared to the uninstrumented binaries, TCS, Slim and Incremental incurred 0.6%, 0.5%, and 0.4% of size increase, respectively.

2) Efficiency of HPAC: To evaluate the run-time overhead of our system, we ran our system on both SPEC CPU2006 Integer benchmarks and a set of real-world service programs.

SPEC CPU2006. The speed overhead incurred by HPAC can be divided into four parts: (1) overhead due to instrumentation, which has been presented above; (2) overhead due to interposition of heap memory allocation calls; (3) overhead due to maintaining the meta data of each buffer (such that our system does not rely on the internal details of the underlying allocator); (4) overhead due to patch deployment, which causes the security measures to be applied to vulnerable buffers.

In order to measure the overhead incurred due to patch deployment, we select a set of allocation-time CCIDs (Calling-Context IDs) as hypothesized vulnerable ones as follows. First, for each benchmark program, we rank all of its allocation-time CCIDs according to their frequencies during the profiling execution (that is, how many heap buffers have been allocated under that calling context). Next, we pick the CCIDs with median frequencies as the hypothesized vulnerable ones. Finally, we regard the heap buffers with those allocation-time CCIDs as ones vulnerable to overflows (the other two vulnerability types are much less expensive to treat), and generate corresponding patches for them.

Figure 8 shows the measurement results. The overhead due to interposition is 1.9%, and the overhead for maintaining the meta data for buffers (plus the interposition overhead) is 4.3%. Note that this part of overhead can be easily eliminated if our system is integrated with the underlying heap allocator. When one patch is installed, the overhead becomes 4.7%, only 0.4% of overhead increase. The total overhead is 5.2% when five patches are installed. One outlier is 400.perlbench, which has the most intensive heap allocations. Table IV records the heap allocation statistics for each SPEC CPU2006 benchmark.

We also measured the memory consumption of benchmark programs with and without our system deployed. To perform the experiments we used a script that can compute the memory overhead in terms of the average Resident Set Size (RSS) for the benchmark programs. That script reads the VmRSS field of /proc/[pid]/status. The sampling rate of that script is 30 times per second; then, the mean of the reading is calculated. Figure 9 shows the memory consumption overhead normalized over native program execution, and the average memory overhead is only 4.3%.

Service Programs. We also evaluated our system on two popular service programs: Nginx and MySQL. We used Nginx 1.2, and measured the throughput overhead by sending requests using Apache Benchmark. Different numbers of concurrent requests were used, and the throughput was compared with that of native execution. The average throughput overhead is only 4.25%.

MySQL 5.5.9 was used, and we applied the built-in test script to measuring the throughput overhead. There was no observable throughput overhead. The memory overhead in both cases was negligible. Note that the memory overhead is proportional with the number of live heap buffers.

Summary. The evaluation shows that HPAC is not only effective but also very efficient. Its most optimized calling context encoding incurs only 0.4% of slowdown, a 6 times of speed boost compared to the original encoding algorithm. The 4.3% speed slowdown, due to allocation/deallocation interposition and metadata maintaining, can be largely eliminated if our system is integrated with the underlying heap allocator. Installing a single hypothesized heap patch only incurs another 0.4% of speed slowdown. It is noticeable the throughput overhead on
real-world service programs is very low or negligible.

Unlike many heap protection systems that typically incur a very high overhead (e.g., 2.5X of slowdown using MemorySanitizer [22] that detects deterministic uninitialized read detection only and 20% of slowdown using Dieharder [15] that provides probabilistic protection) and/or handles only one heap vulnerability type (e.g., HeapTherapy [21]), HPAC handles multiple frequently exploited heap vulnerability types with high efficiency. The patch generation is precise and automatic, and the patch deployment does not require manual intervention and does not modify the program code, guaranteeing that no new bugs are generated.

IX. DISCUSSION AND FUTURE WORK

A limitation is that HPAC can only handle buffer overflows due to continual write or read operations, which are the main form of buffer overflows. Overflows due to discrete read or write cannot be handled by HPAC. Plus, if an overflow runs over an array which is an internal field of a buffer, HPAC cannot detect it. The limitation is common in many existing countermeasures against buffer overflows, such as AddressSanitizer [10] and Exterminator [53].

It may happen that a heap vulnerability can be exploited with multiple CCIDs, and thus the attacker may develop different attack input to exploit buffers with new allocation calling contexts. However, whenever the attack exploits a buffer allocated in a new calling context, our system simply treats it as a new vulnerability and starts another defense cycle. More importantly, based on our evaluation and previous researches on context-sensitive defenses [21], [40], [48], [53], this is rare.

We do not claim that HPAC is to replace existing patching system. It is to complement conventional patching by providing immediate and bug-free protection when the fresh patches are still not mature and may need more time for testing.

For programs that have a large memory profile, to analyze the use-after-free attack, the memory quota for the FIFO queue of freed blocks may be exceeded. In this case, we can replay attacks in multiple executions; specifically, we divide the whole space of CCIDs into \( N \) subspaces, and each of the \( N \) executions defers the deallocation of buffers that have the allocation-time CCIDs in one of the subspaces. Now, each execution is expected to consume \( 1/N \) of the memory.

Some works aim to discover more zero-day heap vulnerabilities before software release [7]–[9], while other works try to pinpoint memory defects by analyzing the core dumps [60], [61]. How to combine them with our system for proactive patch generation and patch generation without attack replay is an interesting problem; we will explore it as our future work.

X. CONCLUSIONS

We have combined heavyweight offline attack analysis, lightweight online defense generation, and program instrumentation to build a new heap memory defense system HPAC. It has overcome the challenge of generating online defenses for handling a variety of vulnerabilities, and even combo vulnerabilities (a buffer that can be exploited by different types of attacks). The task is further complicated when buffer metadata maintenance is transparent to the underlying heap allocators, and does not assume any special custom allocators.

The new heap defense system has many prominent advantages: (1) patch generation without manual efforts, (2) codeless patching, (3) versatile handling of heap buffer overwrite, overread, use after free, and uninitialized read, (4) imposing a very small overhead, and (5) no dependency on specific allocators. In addition, we have proposed targeted calling context encoding, which should interest researchers applying or building calling context encoding techniques. The evaluation
