Security and Cryptography 2021

IV. Cache Attacks

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Cache attacks against a process

- side channel attack via measuring time
- similar mechanism as used for Meltdown: detecting cache misses indicates some particular execution pattern

Example: "Cache Missing for fun and profit" by Colin Percival

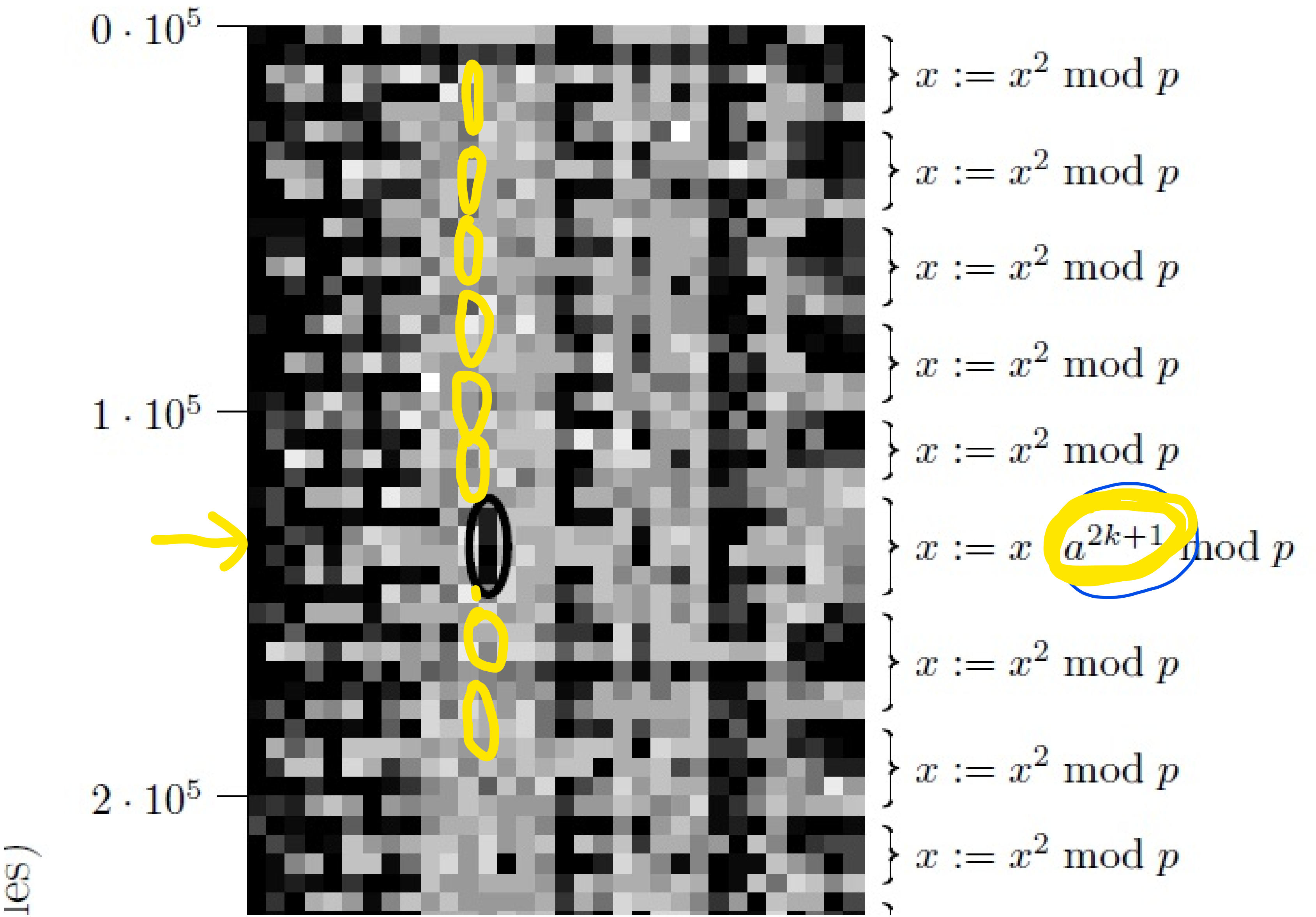
goal: find the RSA private key from OpenSSL executen on Pentium4 (original attack)

practical issues about cache:

- if there is a victim thread and a spy thread then in the time between switching victim to spy the whole L1 can be evicted anyway as it is small
- L1: is very fast, time differences between a hit and miss and fetching from L2 are not big, problematic time measuring with rdtsc by the spy thread
- instructions are not loaded into L1 as to L2, no noise of this kind in L1
- problems with hardware prefetcher: if a few cache misses occurs on subsequent addresses then a few cache line fetched "just for the case" – so the spy process inspecting cache misses must "jump " between addreses
- TLB **(translation lookaside buffer)** misses influence time as well, TLB does not cover whole L2

OpenSSL RSA implementation

- Chinese Remainder Theorem used:
  - instead of computing $a^d \bmod n$, where $n = p \cdot q$
  - one computes $a^d \bmod p$ and $a^d \bmod q$ and combines the results with CRT
  - so: smaller numbers in game
- sliding window exponentiation method
  - precomputed values: $a^3, a^5, \ldots, a^{31} \bmod p$
  - "square and multiply" method: a series of squarings $x := x^2 \bmod p$, and multiplications $x := x \cdot a^{2k+1}$
  - squaring and multiplication use different algorithms with different "footprints" left in the cache
  - footprint also indicates approximately k from $x := x \cdot a^{2k+1}$



$0 \cdot 10^5$
$1 \cdot 10^5$
$2 \cdot 10^5$
$x := x^2 \bmod p$
$x := x^2 \bmod p$
$x := x^2 \bmod p$
$x := x^2 \bmod p$
$x := x^2 \bmod p$
$x := x\ a^{2k+1} \bmod p$
$x := x^2 \bmod p$
$x := x^2 \bmod p$
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Factorization of RSA number n when some bits of p and q are known

starting from $j=1$ to $\log n$ find candidates $p \bmod 2^j$, $q \bmod 2^j$ such that

$$d \cdot e = 1 \bmod (p-1)\cdot(q-1)$$

$$d \cdot e - 1 \bmod 2^j = (p-1)\cdot(q-1) \bmod 2^j$$

$$d \cdot e - 1 \bmod 2^j = (p-1) \bmod 2^j \cdot (q-1) \bmod 2^j \bmod 2^j$$

when we increase j we can prune the some solutions – those that have bits different from the ones already known

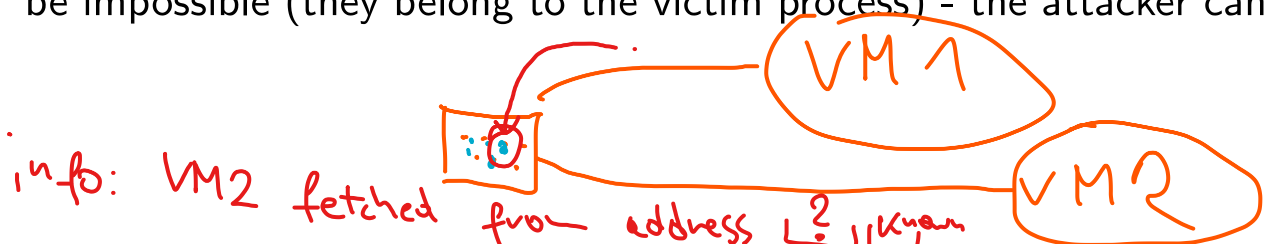
Remark

- libraries often guard against such problems – no subroutines with variable time
- .. but frequently not the case:
  - if the public key not stored but only encrypted secret key, then public key recomputed (ECDSA)
  - computation must be based on plaintext secret key exponentiation

    so: a potential point of leakage via cache timings if sliding window used

Secure processing in a Data Center

- multiprocess architectures, with strict separation between processes offered by the system: hypervisor and virtualization, sandboxing, ...
- an attacker process tries to get secrets from victim processes without having any priviledges
  - theoretically virtualization solves the problem
- despite separation protection the processes share cache
- there is a strict control over the cache content but **cache hits and cache misses** might be detected by **timing for the attacker's process** (and not of the victim process)
- the timing for cache access should somehow depend on the sensitive information to be retreived
- difficulty: other than in the classical cryptanalysis – access to plaintext or ciphertext might be impossible (they belong to the victim process) - the attacker can only predict something



CASE STUDY: AES encryption

AES software implementation:

- particularly vulnerable because of its design
- AES defined in algebraic terms, but lookup table is typically faster
- there are arguments against algebraic implementations as the execution time may provide a side channel
- key expansion: round zero: simply the key bytes directly, other rounds: key expansion reversable (details irrelevant for the attack)
- fast implementation based on lookup tables T_0, T_1, T_2, T_3 and $T_0^{(10)}, T_1^{(10)}, T_2^{(10)}, T_3^{(10)}$ for the last round (with no `MixColumns`)

- round operation

$$\left(x_0^{(r+1)}, x_1^{(r+1)}, x_2^{(r+1)}, x_3^{(r+1)}\right) := T_0(x_0^r) \oplus T_1(x_5^r) \oplus T_2(x_{10}^r) \oplus T_3(x_{15}^r) \oplus K_0^{(r+1)}$$

$$\left(x_4^{(r+1)}, x_5^{(r+1)}, x_6^{(r+1)}, x_7^{(r+1)}\right) := T_0(x_4^r) \oplus T_1(x_9^r) \oplus T_2(x_{14}^r) \oplus T_3(x_3^r) \oplus K_1^{(r+1)}$$

$$\left(x_8^{(r+1)}, x_9^{(r+1)}, x_{10}^{(r+1)}, x_{11}^{(r+1)}\right) := T_0(x_8^r) \oplus T_1(x_{13}^r) \oplus T_2(x_2^r) \oplus T_3(x_7^r) \oplus K_2^{(r+1)}$$

$$\left(x_{12}^{(r+1)}, x_{13}^{(r+1)}, x_{14}^{(r+1)}, x_{15}^{(r+1)}\right) := T_0(x_{12}^r) \oplus T_1(x_1^r) \oplus T_2(x_6^r) \oplus T_3(x_{11}^r) \oplus K_3^{(r+1)}$$

attack notation:

– $\delta = B\,/\,\text{entrysize of lookup table}$, typically: entrysize=4bytes, $\delta = 16$, (so δ entries of a lookup table are within the same cache line – this is a complication for the attack!)

– for a byte y let $\langle y \rangle = \lfloor y/\delta \rfloor$, it indicates a memory block of y in T_l

– if $\langle y \rangle = \langle z \rangle$, then x and y correspond to requests to the same memory block of the lookup table and therefore to the same cache line

– $Q_k(p, l, y) = 1$ iff AES encryption of plaintext p under key K accesses memory block of index y in T_l at least once in 10 rounds

– $M_k(p, l, y) =$ measurement, its expected value is bigger when $Q_k(p, l, y) = 1$ then if when $Q_k(p, l, y) = 0$

“synchronous attack”

- **plaintext random but known**, corresponds to the situation where one can trigger encryption (e.g. VPN with unknown key, dm-crypt of Linux)
- phase 1: measurements, phase 2: analysis
- from experiments: AES key recovered using 65 ms of measurements (800 writes) and 3 sec analysis

attack on round 1:

i accessed indices for lookup tables are simply $x_i^{(0)} = p_i \oplus k_i$ for $i = 0, \ldots, 15$

ii goal: find information $\langle k_i \rangle$ of k_i – one cannot derive information on lsb; candidates for k_i are denoted by $\bar{k_i}$

iii if $\langle k_i \rangle = \langle \bar{k_i} \rangle$ and $\langle y \rangle = \langle p_i \oplus \bar{k_i} \rangle$, then $Q_k(p, l, y) = 1$ for the lookup $T_l(x_i^{(0)})$

iv if $\langle k_i \rangle \neq \langle \bar{k_i} \rangle$, then there is no lookup in block y for T_l during the first round, **but**

- there are $4 \cdot 9 - 1 = 35$ other accesses affected by other plaintext bits during the entire encryption (4 per round, 9 rounds in total as the last round uses different look-up tables)
- probability that none of them accesses block y for T_l is

  $\left(1 - \frac{\delta}{256}\right)^{35} \approx 0.104$ for $\delta = 16$

v few dozens of samples required to find a right candidate for $\langle k_i \rangle$

vi together we determine $\log(256/\delta) = 4$ bits of each byte of the key

vii no more possible for the first round, still 64 key bits to be found, so one cannot do the rest with a brute force

viii in reality more samples needed due to noise in measurements $M_k(p, l, y)$ and not $Q_k(p, l, y)$

attack on round 2: the goal is to find the still unknown key bits

i we exploit equations derived from the Rijndeal specification:

$$x_2^{(1)} = s(p_0 \oplus k_0) \oplus s(p_5 \oplus k_5) \oplus 2 \bullet s(p_{10} \oplus k_{10}) \oplus 3 \bullet s(p_{15} \oplus k_{15}) \oplus s(k_{15}) \oplus k_2$$

$$x_5^{(1)} = s(p_4 \oplus k_4) \oplus 2 \bullet s(p_9 \oplus k_9) \oplus 3 \bullet s(p_{14} \oplus k_{14}) \oplus s(p_3 \oplus k_3) \oplus s(k_{14}) \oplus k_1 \oplus k_5$$

$$x_8^{(1)} = \ldots.$$

$$x_{15}^{(1)} = \ldots.$$

where s stands for the Rijndael Sbox, and $\bullet$ means multiplication in the field with 256 elements

ii lookup for $T_2\big(x_2^{(1)}\big)$:

- $\langle k_0 \rangle, \langle k_5 \rangle, \langle k_{10} \rangle, \langle k_{15} \rangle, \langle k_2 \rangle$ already known
- low level bits of $\langle k_2 \rangle$ influence only low bits of $x_2^{(1)}$ so not important for cache access pattern
- the upper bits of $x_2^{(1)}$ can be determined after guessing low bits of k_0, k_5, k_{10}, k_{15}: there are δ^4 possibilities ($=16^4$)
- a correct guess yields a lookup in the right place

– an incorrect guess: some $k_i \neq \bar{k_i}$ so

$$x_2^{(1)} \oplus \bar{x}_2^{(1)} = c_i \bullet s(p_i \oplus k_i) \oplus c_i \bullet s(p_i \oplus \bar{k}_i) \oplus \ldots$$

where ... depends on different random plaintext bits and therefore random

– differential properties of AES studied for AES competition:

$$\Pr[\, c_i \bullet s(p_i \oplus k_i) \oplus c_i \bullet s(p_i \oplus \bar{k}_i) \neq z] > 1 - \left(1 - \frac{\delta}{256}\right)^3$$

so the false positive for lookup in T_2 at a given block:

- $\left(1 - \frac{\delta}{256}\right)^3$ for computing $T_2(x_2^{(1)})$
- $\left(1 - \frac{\delta}{256}\right)$ for computing each of the remaining invocations of T_2
- together no access with pbb about $\left(1 - \frac{\delta}{256}\right)^{38}$

– this yields about 2056 samples necessary to eliminate all wrong candidates

– it has to repeated 3 more times to get other nibbles of key bytes

iii optimization: guess $\Delta = k_i \oplus k_j$ and take $p_i \oplus p_j = \Delta$, then i.e. $s(p_0 \oplus k_0) \oplus s(p_5 \oplus k_5)$ cancels out and we have to guess less bits (4 instead of 8)

- **similar attack: last round** - created ciphertext must be known to the attacker, otherwise similar. Subkey from the last round learnt, but key schedule is reversible

- **cache measurement strategy: Evict+Time**
  - i procedure:
    1. trigger encryption of a plaintext p
    2. **evict:** access memory addresses so that **one cache set overwritten** completely
    3. trigger encryption of the plaintext p
  - ii in the evicted cache set one cache line from T_l **is missing**
  - iii measure time: if long, then cache miss and the encryption refers to evicted δ positions from the lookup table
  - iv practical problem: triggering may invoke other activities and timing is not precise

- **measurement: Prime+Probe**
    - i procedure
        1. **prime:** overwrite entire cache by reading A: a contiguous memory of the size of the cache
        2. trigger an encryption of p – it results in **eviction** at places where lookup has occurred
        3. **probe:** read memory addresses of A and **detect which locations have been evicted**
    - ii easier: probe timing suffices to check, if encryption used a given cache set
- **complications in practice:**
    - i **address of lookup tables in the memory** - how they are loaded to the cache remains unknown – offset can be found by considering all offsets and then statistics for each offset (experiments show good results even in a noisy environment)
    - ii **hardware prefetcher** may disturb the effects. Solution: read and write the addresses of A according to a pseudorandom permutation
- **practical experiments:** e.g. Athlon 64, no knowledge of adresses mapping, 8000 encryptions with Prime & Probe

  Linux dm-crypt (disk, filesystem, file encryption): with knowledge of addressing, 800 encryptions (65 ms), 3 seconds analysis, full AES key

extensions of the attack:

- on some platforms timing shows also position of the cache line (better resolution for one-round attack)
- remote attacks (VPN, IPSec): with requests that trigger immediate response (situation yet unclear about practicality)

"asynchronous attrack" on round 1

- no knowledge of plaintext, no knowledge of ciphertext
- based on frequency F of bytes in e.g. English texts, frequency score for each of $\frac{256}{\delta}$ blocks of length δ
- F is nonuniform: most bytes have high nibble = 6 (lowercase characters "a" through "o")
- find j such that j is particularly frequent indicates $j = 6 \oplus \langle k_i \rangle$ and shows $\langle k_i \rangle$
- complication: this frequency concerns at the same time k_0, k_5, k_{10}, k_{15} affecting T_0 so we learn 4 nibbles but not their actual allocation to k_0, k_5, k_{10}, k_{15}
- the number of bits learnt is roughly: $4 \cdot (4 \cdot 4 - \log 4!) \approx 4 \cdot (16 - 3.17) \approx$ 51 bits
- experiment: OpenSSL, measurements 1 minute, 45.27 info bits o on the 128-bit key gathered

Bernstein's attack

– an alternative way of computing AES, algorithm applied in OpenSSL:

→ two constant 256-byte tables: S and S'

→ expanded to 1024-byte tables T_0, T_1, T_2, T_3

$$T_0[b] = (S'[b], S[b], S[b], S[b] \oplus S'[b])$$

$$T_1[b] = (S[b] \oplus S'[b], S'[b], S[b], S[b])$$

$$\ldots.$$

→ AES works with 16-byte arrrays x and y, where x initialized with the key k, y initialized with $n \oplus k$, where n is the plaintext

$\rightarrow$ AES computation is modifications of x and y:

i x viewed as (x_0, x_1, x_2, x_3) (4 bytes parts)

ii $e := (S[x_3(1) \oplus 1], S[x_3(2)], S[x_3(3)], S[x_3(0)])$

iii replace (x_0, x_1, x_2, x_3) with $(e \oplus x_0, e \oplus x_0 \oplus x_1, e \oplus x_0 \oplus x_1 \oplus x_2, e \oplus x_1 \oplus x_2 \oplus x_3)$

iv replace $y = (y_0, y_1, y_2, y_3)$ with

$$(T_0[y_0[0]] \oplus T_1[y_1[1]] \oplus T_2[y_2[2]] \oplus T_3[y_3[3]] \oplus x_0,$$

$$(T_0[y_1[0]] \oplus T_1[y_2[1]] \oplus T_2[y_3[2]] \oplus T_3[y_0[3]] \oplus x_1,$$

$$(T_0[y_2[0]] \oplus T_1[y_3[1]] \oplus T_2[y_0[2]] \oplus T_3[y_1[3]] \oplus x_2,$$

$$(T_0[y_3[0]] \oplus T_1[y_0[1]] \oplus T_2[y_1[2]] \oplus T_3[y_2[3]] \oplus x_3$$

v 2nd round uses $\oplus 2$ instead of $\oplus 1$ for x, otherwise the same. Similar changes corresponding to rounds up to 9

vi in round 10 use $S[], S[], S[], S[]$ instead of $T's$

vii y is the final output

it is embarassing how simple the attack is:

→ it has been checked in practice that execution depends on $k[i] \oplus p[i]$ - which is a position in the table:

- try many plaintexts p
- collect statistics for each byte for $p[i]$
- the maximum occurs for z
- the maximum corresponds to a fixed value for $k[i] \oplus p[13]$, say c
- compute $k[13] = c \oplus z$

→ for different bytes different statistics observed: for some t a few values $k[t] \oplus \mathrm{plaintext}[t]$, where substantially higher time observed

→ statistic gathered, different packet lengths

→ finally brute force checking all possibilites, nonce encrypted with the server key

Countermeasures

- ''**no reliable and practical countermeasure**'' so far
- implementation based on **no-lookup** but algebraic algorithm (slow!!!) or bitslice implementation (sometimes possible and nearly as efficient as lookup)
- **alternative lookup tables:** if smaller then smaller leakage (but easier cryptanalysis for small Sboxes)
- **data-independent access to memory** blocks - every lookup causes a redundant read in all memory blocks, generally: oblivious computation possible theoretically, but overhead makes it rather useless
- **masking operations:** $\approx$"we are not aware of any method that helps to resist our attack"
- **cache state normalization:** load all lookup tables - equires deep changes in OS and reduces efficiency, even then LRU cache policy may leak information which part has been used!
- **process blocking:** again, deep changes in OS
- **disable cache sharing**: deep degradation of performance

- **"no-fill" mode** during crypto operations:
  - preload lookup tables
  - activate "no-fill"
  - crypto operation
  - deactivate "no-fill"

  the first two steps are critical and no other process is allowed to run

  possible only in priviledged mode, cost of operation prohibitive
- **dynamic table storage:** e.g. many copies of each table, or permute tables

  details architecture dependent and might be costly
- **hiding timing information:** adding random values to timing makes the statistical analysis harder but still feasible
- **protect some rounds** (the first 2 and the last one) with any mean – but may be there are other attack techniques...
- **cryptographic services at system level**: good but unflexible
- **sensitive status for user processes:** erasing all data when interrupt
- **specialized hardware support:** crypto co-processor seems to be the best choice

  but the problem is not limited to AES or crypto – many sensitive data operations are not cryptographic and a coprocessor does not help