Computational Algebra
-
Transcript

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Contents

1 Integer Arithmetic

Topics:

- Addition and Multiplication
- GCD computation
- Primality testing
- Factorization

1.1 Addition and Multiplication

Agreement:

- $a, x \in \mathbb{N}$ represented as $x = \sum_{i=0}^{n-1} a_i \cdot B^i \quad B \in \mathbb{N}_{>1}$ fixed Base ($a_i \in \{0,, B-1\}$)
- if $x \neq 0$, assume $a_{n-1} \neq 0$ then define:
  length of $x := l(x) = n\;$ = number of digits = $\lfloor \log_B(x) \rfloor + 1$
  (mnemonic: $\log_B(B) + 1 = 2$)
- $l(0) = 1$
  (Amount of memory required to store $x = 0$)
- $l(x) := l(|x|)$
- for $x \in \mathbb{Z}$ represent if as $x = sgn(x) * |x|$

1.1.1 Algorithm 1 (Simple addition)

input : $x = \sum_{i=0}^{n-1} a_i \cdot B^i, \quad y = \sum_{i=0}^{n-1} b_i \cdot B^i, \quad x, y \in \mathbb{N}$

output: $x + y = \sum_{i=0}^{n} c_i \cdot B^i$

(1) $\sigma = 0$

(2) for $i = 0, ..., (n-1)$:

(3) set $c_i := a_i + b_i + \sigma_i$ and $\sigma := 0$

(4) if $(c_i \geq B)$

(5) set $c_i = c_i - B$

(6) set $\sigma = 1$

(7) set $c_n = \sigma$

If $B = 2$ then (3) - (6) can be realized by logic gates:

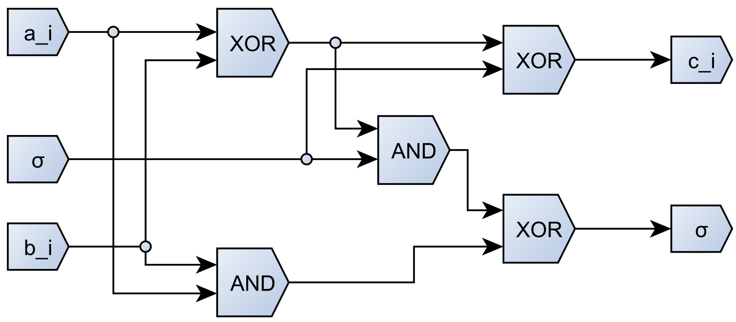

Figure 1: Logic circuit for addition

1.1.2 Definition 2 (Bit-Operation)

A bit operation is an operation that can be performed by a logic gate or by searching or writing a bit from / into memory.

1.1.3 Definition 3 (Big O)

Let M be a set (usually $M = \mathbb{N}$), $\quad f, g : M \mapsto \mathbb{R}_{>}0$
we write $f \in O(g) \quad$ if $\quad \exists c \in \mathbb{R} : f(x) \leq c \cdot g(x) \forall x \in M$

1.1.4 Theorem 4 (Lower bound for addition)

Let $f : \mathbb{N} \mapsto \mathbb{R}, \quad n \mapsto$ maximal number of bit operations required by Algorithm 1 to add $x_y \in \mathbb{N}$ with $l(x), l(y) \leq n$
Let $g = id_{\mathbb{N}}$ Then $f \in O(g)$
We say Algorithm 1 requires $O(n)$ bit operations for adding two numbers of length $\leq n$.
$\Rightarrow$ "linear complexity"
Set $M := \{$Set of all algorithms for addition in $\mathbb{N}\}$
For $A \in M$ define $f_A : \mathbb{N} \mapsto \mathbb{R}$ as above.
We would like to find $f_{odd} : \mathbb{N} \mapsto \mathbb{R}, \quad n \mapsto \inf\{f_A(n) | A \in M\}$
Since one needs to read x, y (and write the result) we can not do any better than linear complexity for addition.
Subtraction
let x, y as Algorithm 1, $x \geq y$
For $\bar{y} := \sum_{i=0}^{n-1} (B - 1 - b_i) B^i \qquad$ (digitwise / bitwise complement)
$\Rightarrow x + \bar{y} = x - y + B^n - 1$
$\Rightarrow x - y = x + \bar{y} + 1 - B^n \qquad$ (initially set $\sigma = 1$)
Conclusion: Addition and Subtraction have cost $O(n)$

1.1.5 Algorithm 5 (Multiplication by "grid method")

input : $x = \sum_{i=0}^{n-1} a_i \cdot 2^i, \quad y = \sum_{i=0}^{m-1} b_i \cdot 2^i$

output: $z = x \cdot y$

(1) $z := 0$

(2) for $i = 0, ..., (n-1)$

(3) if $(a_i \neq 0)$ set $z := z + \sum_{j=1}^{m-1} b_j 2^{i+j}$

1.1.6 Theorem 6 (Runtime of Algorithm 5)

Algorithm 5 requires $O(n * m)$ bit operations.
As of the total input length $n + m$:
$n \cdot m \leq \frac{1}{2}(n+m)^2 \rightarrow O((n+m)^2)$
$\Rightarrow$ Quadratic complexity

Karatsuba-multiplication:
Observation for polynomials:
$a + bx, c + dx$ have $(a+bx)(c+dx) = ac + (ac + db - (a-b)(c-d))x + bdx^2$
The point: only used 3 multiplications instead of 4.
Specialize $x = B$ "large" such that $x = a + bB$ partition into two blocks. Then multiply the blocks by a recursive call.

1.1.7 Algorithm 7 (Karatsuba)

input : $x, y \in \mathbb{N}$

output: $z = x \cdot y$

(1) Choose $k \in \mathbb{N}$ minimal such that $l(x), l(y) \leq 2^k$.
Set $B = 2^{2^{k-1}}$

(2) if $(k = 0)$ return $x \cdot y$ (by bit-operation AND)

(3) write $x = x_0 + x_1 B, \quad y = y_0 + y_1 B$ with $l(x_i), l(y_i) \leq 2^{k-1}$

(4) compute $x_0 \cdot y_0, \; x_1 \cdot y_1, \; (x_0 - x_1) \cdot (y_0 - y_1)$ by a recursive call

(5) return $z = x_0 y_0 + (x_0 y_0 + x_1 y_1 - (x_0 - x_1)(y_0 - y_1))B + x_1 y_1 B^2$

1.1.8 Theorem 8 (Runtime of Algorithm 7)

For multiplying two numbers of length $\leq n$ Algorithm 7 requires $O(n^{\log_2 3}) \approx O(n^{1.59})$ bit operations.
Proof:
Set $\Theta(k) :=$ maximal numbers of bit operations for $l(x), l(y) \leq 2^k$
We have for $k > 0 : \Theta(k) \leq 3\Theta \underbrace{(k-1)}_{\text{recursive calls}} + c \underbrace{2^k}_{\text{additions}}$ with (c some constant)
Claim: $\Theta(k) \leq 3^k + 2c(3^k - 2^k)$
Proof by Induction on k:
$k = 0 : \Theta(k) = 1$
$k - 1 \to k = \Theta(k) = 3\Theta(k-1) + c2^{k-1}$
$\leq 3(3^{k-1} + 2c(3^{k-r} - 2^{k-1})) + c2^k$
$= 3^k + 2c(3^k - 2^k)$
So $\Theta(k) \leq (2c+1)3^k$
Now $l(x) \leq n$ hence $2^{k-1} < n$ by minimality of k
So $k - 1 < \log_2 n$
$\Rightarrow \Theta(k) \leq 3(2c+1)3^{\log_2(n)}$
$= 3(2c+1)2^{\log_2(3)\log_2(n)}$
$= 3(2c+1)n^{\log_2(3)}$ $\square$
One can modify the terminal condition of Karatsuba to switch to Grid-Multiplication, which is faster for small numbers.

Fast-Fourier Transform
Reminder: For a function $f : \mathbb{R} \mapsto \mathbb{C}$ define:
$\hat{f} : \mathbb{R} \mapsto \mathbb{C}$ by
$\hat{f}(\omega) = \int\limits_{\mathbb{R}} f(t)e^{-i\omega t}dt$ (if it exists)
Think of ω as frequency.

Definition (Convolution)
Let $f, g : \mathbb{R} \mapsto \mathbb{C}$
$(f * g)(x) = \int\limits_{\mathbb{R}} f(t)g(x-t)dt$
Convolution is analogous to polynomial multiplication **Formula:** $\underbrace{(f * g)}_{\text{(Cauchy formula)}} = \hat{f} \cdot \hat{g}$
For a function $M \mapsto C$ with $|M| < \infty$ we need the discrete Fourier transform (DFT)

1.1.9 Definition 9 (Root of unity)

Let R be a commutative ring with 1. An element $\mu \in R$ is called an $n-th$ root of unity (= root of 1) if $\mu^n = 1$.
It is called primitive if $\mu^i \neq 1$ for $(0 < i < n)$ i.e. $ord(\mu) = n$
let μ be a primitive n-th root of 1 (e.g. $e^{2\pi \frac{i}{n}} \quad \in \mathbb{C}$)
Then the map $DFT_\mu : R^n \mapsto R^n$
$(\hat{a}_0, ..., \hat{a}_n) \mapsto (\hat{a}_0, ..., \hat{a}_n) \quad$ with $\hat{a}_i = \sum\limits_{j=0}^{n-1} \mu^{ij} a_j$
is called discrete Fourier transformation
For polynomials:
$DFT_\mu : R[x] \mapsto R^n$
$\qquad f \ \mapsto (f(\mu^0), ..., f(\mu^{n-1})$
Convolution rule: (from $f(\mu^i)g(\mu^i) = (f * g)(\mu^i)$)
$DFT_\mu(f * g) = DFT_\mu(f) \cdot DFT_\mu(g) \qquad$ (component wise product)

Addition of two polynomials in R[x] of $deg(n)$ require $O(n)$ ring operations.
Multiplication require $O(n^l)$.
With Karatsuba have $O(n^{\log_2(3)})$ ring operations.
Cost $DFT_\mu(f) \cdot DFT_\mu(g) : O(n)$ ring operations $\quad$ (with μ as 2n-th root of 1)
Want: Cheap way of doing DFT and back-transformation.

1.1.10 Algorithm 10 (Fast Fourier transformation FFT)

input : $f \in R[x], \quad \mu \in R$ primitive 2^k-th root of 1, such that $\mu^{2^{k-1}} = -1$

output: $DFT_\mu(f)$

(1) Write $f(x) = g(x^2) + xh(x^2) \quad$ with $f, g, h \in R[x]$

(2) if $(k = 1) \qquad //(\Rightarrow \mu = 1)$
return $DFT_\mu(f) = (g(1) + h(1), g(1) - h(1))$

(3) Recursive call: compute $DFT_{\mu^2}(g) = \hat{g}, DFT_{\mu^2}(h) = \hat{h} \quad \in R^{2^{k-1}}$

(4) return $DFT_\mu(f) = (\hat{f}_0, ..., \hat{f}_{2^k-1}) \quad$ with $\hat{f}_i = \hat{g}_i + \mu\hat{h}_i$
where $\hat{g}_i = \hat{g}_{i-2^{k-1}} \quad$ for $i \geq 2^{k-1}$

Note: Components of $\hat{g}$ and $\hat{h}$ are:
$\hat{g} = g(\mu^{2i}), \quad \hat{h}_i = h(\mu^{2i}) \quad$ so
$\hat{f}_i := f(\mu^i) = \hat{g}_i(\mu^{2i}) + \mu h(\mu^{2i}) = \hat{g}_i + \mu\hat{h}_i$

Convention: $\lg(x) = \log_2(x)$

1.1.11 Theorem 11 (Runtime of Algorithm 10)

Let $n = 2^k, \quad f \in R[x] \quad$ with $deg(\psi) < n$
Then Algorithm 10 requires $O(n \cdot \lg(n))$ ring operations.
Better than $O(n^{1+\epsilon}), \quad \forall \epsilon > 0$!
Proof:
Set $\Theta(k)$ = max number of ring operations required. By counting obtain for $k > 1$:
$\Theta(k) \leq 2\Theta(k-1) + \underbrace{(\text{compute } \mu^i (i \leq 2^{k-1}))}_{2^{k-1}} + \underbrace{(\mu^i \hat{k}_i)}_{2^{k-1}} + \underbrace{(\text{sums and differences})}_{2^k}$
$= 2\Theta(k-1) + 2^{k+1}$
Claim: $\Theta(k) \leq (2k-1)2^k$
$k = 1 : f = a_0 + a_1 \cdot x \quad DFT_\mu(f) = (a_0 + a_1 \cdot a_0 - a_1) \Rightarrow \Theta(a) = 2$
$k-1 \to k : \Theta(k) \leq 2 \cdot \Theta(k-1) + 2^{k+1} \leq 2 \cdot (2k-3) \cdot 2^{k-1} + 2^{k+1} = (2k-1) \cdot 2^k$
since $k = \lg(n)$ obtain $O(k) \leq (2 \cdot \lg(n) - 1) \cdot n \in O(n \cdot \lg(n)) \qquad \square$

Back-transformation?

1.1.12 Definition 12 (Good root of unity)

A primitive n-th root of unity is called good (caveat: this is ad-hoc terminology) if:
$\sum_{j=0}^{n-1} \mu^{ij} = 0 \quad$ for $(0 < i < n)$
example:

(1) $\mu = e^{2\pi \frac{i}{n}}$ is a good primitive root of unity

(2) $R = \mathbb{Z}/(8), \quad \mu = \bar{3} \Rightarrow \mu \cdot B$ is primitive 2^{nd} root of unity
But $\bar{B}^0 + \bar{3}^1 = \bar{u} \neq \bar{0} \qquad$ so μ is not good.

1.1.13 Proposition 13 ($DFT_{\mu^{-1}}$)

let $\mu \in R$ be a good root of 1
$(a) = (a_0, ..., a_{n-1}) \in R^n \Rightarrow DFT_\mu^{-1}(DFT_\mu(a)) = n \cdot (a) \quad$ where $n = 1 + ... + 1 \in R$
Proof:
$DFT_\mu(a) = (\hat{a}) = (\hat{a}_0, ..., \hat{a}_{n-1})$
with $\hat{a}_j = \sum_{k=0}^{n-1} \mu^{jk} a_k$
$DFT_{\mu^{-1}}(\hat{a}) = (\hat{\hat{a}}_0, ..., \hat{\hat{a}}_1)$

$$\text{with } \hat{\hat{a}}_i \sum_{i=0}^{n-1} \mu^{-ij} \sum_{k=0}^{n-i} \mu^{jk} a_k = \sum_{k=0}^{n-1} \left(a_k \cdot \underbrace{\sum_{i=0}^{n-1} \mu^{j(k-i)}}_{=0 \text{ if } n \neq k-i \ \ (\text{i.e.} k=i)} \right) = a_i \cdot n \qquad \square$$

1.1.14 Proposition 14 (Finding good roots of unity)

let $\mu \in R, n \in \mathbb{N}$
Assume:

a) R is an integral Domain and μ is a primitive or n-th root of 1
(Integral Domain: nonzero commutative ring in which the product of two nonzero elements is nonzero)
$\Rightarrow$ Granted by FFT

b) $n = 2^b$, $\mu^{\frac{n}{2}} = -1$, then $h > 0 \wedge char(R) \neq 2$
$\rightarrow \mu$ is a good primitive n-th root of 1 ("root of unity")

Proof:

a) for $0 < i < n$
$$\underbrace{(\mu^i - 1)}_{\neq 0}\underbrace{(\sum_{j=0}^{n-1} \mu^{ij})}_{=0\ *} = \mu^{in} - 1 = 0$$
$\Rightarrow \mu$ is a good root of unity

* Let $0 < i < n$, write $i = 2^{k-s} \cdot r$ with r odd $\wedge\ s > 0$
$$\sum_{j=0}^{2^k-1} \mu^{ij} = \sum_{l=0}^{2^{k-s}-1} \sum_{j=0}^{2^s-1} \mu^{i(l \cdot 2^s + j)}$$
$$\mu^{i \cdot 2^s} = 1$$
$$i \cdot 2^s = 2^{k-s} \sum_{j=0}^{2^s-1} \mu^{ij} = 2^{k-s} \sum_{j=0}^{2^{s-1}-1} (\mu^{ij} + \mu^{i(2^{s-1}+j)})$$
But $\mu^{i \cdot 2^{s-1}} = \mu^{2^{k-s} \cdot r \cdot 2^{s-1}} = \mu^{2^{k-1} \cdot r} = (-1)^r = -1$
So $\sum_{j=0}^{n-1} \mu^{ij} = 0$ $\square$

b) $\mu^n = 1, n = 2^k \Rightarrow ord(\mu) | n \Rightarrow ord(\mu)$ is power of 2

1.1.15 Algorithm 15 (Polynomial multiplication using DFT)

input : $f, g \in R[x]$ with $deg(f) + deg(g) < 2^k =: n$
$\mu \in R$ as a good root of unity; Assume $2 \in R$ is invertible

output: $h = f \cdot g$

(1) compute $\hat{f} = DFT_\mu(f)$, $\hat{g} = DFT_\mu(g)$ with $f, g \in R^n$

(2) compute $\hat{h} = \hat{f} \cdot \hat{g}$

(3) compute $(h_0, ..., h_{n-1}) = DFT_{\mu^{-1}} \hat{h}$ (same as $DFT_\mu(\hat{h})$ but with different order)
= Back-transformation $\cdot 2^k$
set $h = \frac{1}{2^k} \sum_{i=0}^{n-1} h_i x^i$

1.1.16 Theorem 16 (Runtime of Algorithm 15)

Algorithm 15 uses $O(n \cdot \log(n))$ ring operations for polynomials of $deg < n$
Proof:

- Choose k minimal so that $deg(f) \cdot deg(g) < 2^k$
  $\Rightarrow 2^{k-1} \leq 2n \quad \Rightarrow k \leq \log(n) + 2$
- $\underbrace{O(2k \cdot 2^k)}_{\text{Step 1}} + \underbrace{2^k}_{\text{Step 2}} + \underbrace{O(k \cdot 2^k) + 2^k}_{\text{Step 3}} \in O(2k \cdot 2^k) = O(n(g(n))) \qquad \square$

Goal: Multiplication in $\mathbb{N}$ using DFT
Idea: find roots of 1 in a suitable $\mathbb{Z}/(m)$
Choose $m = 2^l + 1, \mu = \bar{2} \in R$

1.1.17 Proposition 17 (Add and mul in $O(l)$)

Let $m = 2^l + 1,\ R = \mathbb{Z}/(m)$
Addition in R and multiplication by $\bar{2}^i \in R$ $(0 \leq i < 2l)$ can be done in $O(l)$ bit operations
Proof:

Let $\bar{x} \in R \quad$ with $0 \leq x \leq 2^l$

- Addition: $x \bar{+} \bar{y}$
  (1) compute $x + y \in \mathbb{N}: \qquad O(l)$
  (2) if $x + y > 2^l + 1 \quad$ subtract $2^l + 1: \qquad O(l)$
- Multiplication by $\bar{2}^i \quad (0 \leq i < l)$
  (1) Bit-shift i Bits to the left by relocating in memory:
  $$\underbrace{O(length(i))}_{\text{compute addr. of new first bit}} + \underbrace{l}_{\text{copying}} = O(\log(l)) + l \in O(l)$$
- Multiplication by $\bar{2}^i \quad (l \leq i < 2l - 1)$
  (1) Multiplication by $\bar{2}^{i-l}: \qquad O(l)$
  (2) take negative $\bar{2}^i \cdot \bar{x} = -\bar{2}^{i-l} \cdot \bar{x}: \qquad O(l)$

1.1.18 Proposition 18 (Sort of summary)

Let $k, r \in \mathbb{N}, \quad r > 0, \quad m = 2^{2^k \cdot r} + 1, \quad R = \mathbb{Z}/(m), \quad \mu = \bar{2}^r \in R$
$\Rightarrow 2 \in R$ is invertible, μ is a good primitive 2^{k+1}-th root of 1
$\Rightarrow \mu^{2^k} = 1$
Proof: $\rightarrow$ from above

1.1.19 Algorithm 19 (Multiplication using FFT)

input : $x, y \in \mathbb{N}$

output: $Z = x \cdot y$

(1) Choose $k \in \mathbb{N}$ minimal such that $l(x), l(y) \leq 2^{2k}$

(2) if $k \leq 3$, compute $z = x \cdot y$ by Algortithm 5

(3) set $B = 2^{2^k}$, $\quad m = 2^{2^k \cdot 4} + 1$, $\quad R = \mathbb{Z}/(m)$, $\quad \mu = \bar{2}^4 \in R$
(⇒ so μ is a good primitive 2^{k+1}-th root of 1)

(4) write $x = \sum\limits_{i=0}^{2^k-1} x_i \cdot B^i$, same for y with $(0 \leq x_i, y_i < B)$
possible since $x, y < 2^{2^{2k}} = 2^{2^k \cdot 2^k} = B^{2^k}$

(5) compute: $\hat{x} = DFT_\mu(\bar{x}_0, ..., \hat{x}_{2^k-1}, \underbrace{0, ..., 0}_{2^k \text{zeros}}) \in R^{2^{k+1}}$
same for y
→ use FFT

(6) compute: $\hat{z} = \hat{x} \cdot \hat{b} \quad \in R^{2^{k+1}}$ (component wise multiplication)
Perform multiplication in R as follows:
Multiply representatives (non negative and $< m$) by recursive call.
Then reduce modulo m by "negative bit shift" (see proof of Proposition 17)

(7) compute: $(\bar{z}_0, ..., \bar{z}_{2^{k+1}}) = \frac{1}{2^{k+1}} DFT_{\mu^{-1}}(\hat{z}) \quad \in R$ with $0 \leq z < m$

(8) set $z := \sum\limits_{j=0}^{2^{k+1}-1} z_i \cdot B^i$

1.1.20 Theorem 20 (Runtime of Algorithm 19)

Algorithm 19 correctly computes $t = x \cdot y$ and requires $O(n \cdot (\log n)^4)$ bit operations for $l(x), l(y) \leq n$

Proof: Correctness

write $x(t) \sum_{i=0}^{2^k-i} x_i t^i \quad \in \mathbb{Z}[t], \quad y(t), \quad \bar{x}(t) \in R[t], \bar{y}(t), \bar{z}(t)$

by Proposition 18 and Proposition 13 we have $\bar{z}(t) = \bar{x}(t) \cdot \bar{y}(t)$

The l-th coefficient of $x(t) \cdot y(t)$ is $0 \leq \sum_{i+j=l} x_i \cdot y_j < 2^k \cdot B^2 = 2^{k+2 \cdot 2^k} \leq 2^{2^{k+2}} < m$

So $z(t) = x(t) \cdot y(t) \Rightarrow z = z(B) = x(B) \cdot y(B) = x \cdot y$ Cost:

Write $\Theta(k) :=$ max number of bit operations

Analyze Steps:

(1) compute max $\{l(x), l(y)\}$: $\quad O(l(n)) = O(k)$

(2) $O(1)$

(3) no bit operations

(4) compute starting addresses of x_i, y_i in memory: $2 * 2^k$ increments of the address: $2 \cdot 2 \cdot 2^k = 2^{k+2}$ bit ops
$\Rightarrow O(2^k)$

(5) By Theorem 11 need $O(2 \cdot 2^{k+1} \cdot (k+1))$ operations in R which are additions and multiplications by powers of $\bar{z}$ costing $O(2^{k+2})$ bit operations.
Total for (5): $O(k \cdot 2^{2 \cdot k})$

(6) 2^{k+1} multiplications of numbers $< m$, i.e. of length $\leq 2^{k+2}$.
So $k' \leq \frac{k+3}{2}$ for k': the "new" k used in the next recursion level.
For $\alpha \in R_{>0}$ define $\Theta(\alpha) := \Theta(\lfloor \alpha \rfloor)$
Total for (6): $2^{k+1}(\Theta(\frac{k+3}{2}) + \underbrace{O(2^{k+2})}_{\text{reduction } mod\ m})$

(7) For $DFT_{\mu^{-1}}(\hat{z}) : O(k \cdot 2^{2 \cdot k})$ as (5) Since $\bar{z}$ is a n root of 1, multiplication by $\bar{2}^{-k-1}$ is multiplication by a positive power of $\bar{2}$, which costs $O(2^{k+2})$
Total for (7): $O(k \cdot 2^{2 \cdot k})$

(8) For $j \leq 2^{k+1}$ have $\sum_{i=0}^{j-1} z_i \cdot B^i \leq (m-1) \sum_{i=0}^{j-1} B^i = (m-1)\frac{B^j-1}{B-1} < 2(m-1)\frac{B^j}{B} = 2^{1+2^{k+2}+(j-1)2^k}$ so the sum has length $(j+3) \cdot 2 + 1$
Adding $z_j \cdot B^j$ to this sum happens at $(j \cdot 2^k)$-th bit and higher $\Rightarrow$ cost is $O(2^k)$
Total for (8): $O(2^{2 \cdot k})$

Grad total: For $k \geq 4$:
$\Theta(k) \leq 2^{k+1} \cdot \Theta(\frac{k+3}{2}) + c \cdot k \cdot 2^{2 \cdot k} \quad$ with c constant
Also for $k \in \mathbb{R}_{\geq 4}$

Define $\Lambda(k) := \frac{\Theta(k)}{2^{2\cdot k}} \Rightarrow \Lambda(k) \leq \frac{2^{k+1}\Theta(\frac{k+3}{2})}{2^{2\cdot k}} + c \cdot k = 16 \cdot \Lambda(\frac{k+3}{2}) + c \cdot k$
Define $\Omega(k) := \Lambda(k+3)$ So for $k \in \mathbb{R}_{\geq}1$
$\Omega(k) \leq 16 \cdot \Lambda(\frac{k}{2}+3) + c \cdot (k+3) = \underbrace{16\Omega(\frac{k}{2}) + c \cdot (k+3)}_{*}$

Claim: For $i \in \mathbb{N}$ with $2^{i-1} \leq k-3$ have:
$\Lambda(k) \leq 16^i \Omega(\frac{k-3}{2^i}) + c \cdot (k+3)(1+8+...+8^{i-1}) + 3 \cdot c \cdot (1+16+...+16^{i-1})$
Proof by induction:
$i = 0 \Lambda(k) = \Omega(k-3)$
$i \to i+1 : \Lambda(k) \leq 16^i \Omega(\frac{k-3}{2^i}) + c\cdot(k-3)(1+...+8^{i-1}) + 3\cdot c\cdot(1+...+16^{i-1}) \leq 2^i \leq k-3 \quad *$
$\leq 16^i(16\Omega(\frac{k-3}{2^i+1})) + c(\frac{k-1}{2^i}+3) + c(k-3)... =$ claimed result
Take $u \in \mathbb{N}$ minimal with $2^u > k-3 \Rightarrow \Omega(\frac{k-3}{2^u}) \leq \Omega(\lfloor \frac{k-3}{2^u} \rfloor) = \Omega(0) =: D$ (constant)
Note: u roughly is recursion depth
Have $2^{u-1} \leq k-3 \xRightarrow[\text{claim}]{} \Lambda(k) \leq 16^u \cdot D + c \cdot \underbrace{(k-3)}_{<2^u} \cdot \frac{8^u-1}{7} + 3c \cdot \frac{16^u-1}{15} \quad \in O(16^u)$

Have $2^{u-1} \leq k-3 \Rightarrow u \leq \lg(k-3)+1$
$\Rightarrow \Lambda(k) \in O(16^{\lg(k-3)}) = O((k-3)^4)$
$\Rightarrow \Theta(k) = 2^{2\cdot k} \cdot \Lambda(k) \in O(2^{2k} \cdot (k-3)^4)$
Have $2^{2(k-1)} < \underbrace{n}_{\max\{l(x)\cdot l(y)\}} \Rightarrow k \leq \frac{\lg(n)}{2}+1$
So $\Theta(k) \in O(n \cdot (\lg(n))^4)$ $\square$

1.1.21 Theorem 21 (Schönhage-Strassen 1971)

Multiplication of integers of length $\leq n$ can be done in $O(n \cdot \lg(n) \cdot \lg(\lg(n)))$ bit operations. Schönhage-Strassen is used for integers of length $\geq$ 100.000.
Asymptotically faster: Fürer's algorithm.

Comments on Bit complexity

1. Memory requirement may explode!
   $\Rightarrow$ No Problem as bit complexity is upper bound for memory requirements, since memory access is included in bit operations
   ($\to$ only store what is calculated)

2. Computation of addresses in memory take time
   $\Rightarrow$ length of addresses $\approx$ lg(memory space) computations of addresses $\approx$ lg(memory space)^2

3. As memory requirement gets larger access times will get longer.
   $\Rightarrow$ transportation time for data $\geq \frac{\text{diameter of physical storage}}{2*\text{speed of light}}$

1.2 Division with remainder, Euclidean algorithm

1.2.1 Algorithm 1 (Division with remainder)

input : $b = \sum_{i=0}^{n-1} b_i 2^i \quad a = \sum_{i=0}^{n+m-1} a_i 2^i \qquad$ with $a_i, b_i, \in 0, 1, \quad b_{n-1} = 1$

output: $r, q \in \mathbb{N} \quad$ such that $a = q \cdot qb + r, \quad 0 \leq r < b$

(1) $r = a_i \quad q = 0$

(2) for $i = m, m-1, ..., 0$ do

(3) $\qquad$ if $r \leq 2^i \cdot b \quad$ then set $r := r - 2^i \cdot b, \quad q = q + 2^i$

1.2.2 Proposition 2 (Runtime of Algorithm 1)

Algorithm 1 is correct and requires $O(n \cdot (m+1))$ bit operations.
Proof:
Always have $a = q \cdot b + r$
Claim:
before setp (3), have $0 \leq 2^{i+1} \cdot b$
$i = m; \quad 0 \leq r = a < 2^{m+n} = 2^{m+1}c \cdot 2^{n-1} \leq 2^{m-1} \cdot b \quad i < m$ By step (3)
So after last passage through the loop $0 \leq r < b$
Running Time: In step(3), have comparison and (possiby) subtraction. Only n bits involved $\Rightarrow O(n)$
Total: $O(b \cdot (m+1))$

Remarks:

(1) Division with remainder can be reduced to multiplication.
Precisely: given an algorithm for multiplication that requires $M(n)$ bit operations, there exists an algorithm for division with remainder that requires $O(M(n))$ bit operations.

(2) Practically relevant:
Jebelean's algorithm (1997): $O(n^{\lg 3})$

(3) Alternatively, may choose $r\mathbb{Z} \quad$ such that $\lfloor \frac{-b}{2} \rfloor < r \leq \lfloor \frac{b}{2} \rfloor$

(4) Algorithm 1 extends to $\mathbb{Z}$.

(5) All Euclidean rings have division with remainder (by definition).
(e.g., $R = K[x] \quad \rightarrow$ polynomial ring over field,
$R = \mathbb{Z}[i] = \{a + bi | a, b \in \mathbb{Z}\} \subseteq \mathbb{C}, \quad i^2 = -1$)

1.2.3 Algorithm 3 (Euclidean algorithm)

input : $a, b \in \mathbb{N}$

output: $\gcd(a, b)$ "greatest common divisor"

(1) set $r_0 := a, \quad r_i := b$

(2) for $i = 1, 2, 3, ...$ perform steps (3) and (4)

(3) if $r_i = 0$ then $\gcd(a, b) = |r_{i-1}|$

(4) Division with remainder: $r_{i-1} = q \cdot r_i + r_{i+1} \qquad r_{i+1} \in \mathbb{Z}$
$|r_{i+1}| \leq \frac{1}{2}|r_i|$

Example:

$$a = 287, \quad b = 126$$

$$287 = 2 \cdot 126 + 35 \tag{1}$$
$$126 = 4 \cdot 35 - 14 \tag{2}$$
$$35 = (-2) \cdot (-14) + 7 \tag{3}$$
$$-14 = (-2) \cdot 7 + 0 \tag{4}$$

$$\text{So: } 7|(-14) \underset{(3)}{\Longrightarrow} 7|35 \underset{(2)}{\Longrightarrow} 7|126 \underset{(1)}{\Longrightarrow} 7|287$$

On the other hand take a common divisor d; $\quad d|287; \quad d|126$
$\underset{(1)}{\Longrightarrow} d|d \underset{(2)}{\Longrightarrow} d|14 \underset{(3)}{\Longrightarrow} d|7$

1.2.4 Theorem 4 (Correctness of Algorithm 3)

Algorithm 3 is correct.
Proof:
Since $r_{i-1} = q \cdot r_i + r_{i+1}$ every integer $x \in \mathbb{Z}$ satisfies the equivalcence $x|r_{i-1}$ and $x|r_i \Leftrightarrow x|r_{i+1}$ and $x|r_i$ so $\gcd(r_{i-1}, r_i) = \gcd(r_i, r_{i+1} = \gcd(a, b)$
when terminating have $\gcd(a, b) = \gcd(r_{i-1}, 0) = |r_{i-1}| \qquad \square$

1.2.5 Theorem 5 (Runtime of Algorithm 3)

Algorithm 3 requires $O(m \cdot n)$ bit operations for $n = l(a), m = l(b)$
Proof:
If $a < b$ than the first passage yields $r_2 = a,\ \ r_1 = b$. Cost: $O(n)$
May assume: $a \geq b$. Write $n_i = l(r_i)$
By Proposition 2 $\exists c$ constant such that the total time is $\leq c \cdot \underbrace{\sum_{i=1}^{k} n_i \cdot (n_{i-1} - n_i + 1)}_{=:\sigma(n_0,\dots,n_k)}$

For $i > 2:\quad n_i = n_{i-1} - 1$
Special Case: $n_i = n_{i-1} - 1\quad$ for $i \geq 2$
$\Rightarrow n_i = n_i - i + 1,\quad n_i = m,\quad k = m + 1$
Obtain $\sigma(n_0, ..., n_k) = m\cdot(n-m+1) + \sum_{i=2}^{m+1}(m-i+1)\cdot 2 = m\cdot n - m^2 + m + m(m-1) = m\cdot n.$
Claim: The special case is the worst (most expensive)!
From any sequence $n_1 > n_2 > ... > n_k$ get to the special case by iteratively inserting numbers in the gaps. Insert s with $n_{j-1} > s > n_j$.
$\sigma(n_0, ..., n_{j-1}, s, n_j, ..., n_k) - \sigma(n_0, ..., n_k) = ... = s + (n_{j-1} - s) \cdot (s - n_j)$
$sp\sigma(n_0, ..., n_k) \leq \sigma(n, m, m-1, ..., 2, 1, 0) = n \cdot m \qquad \square$
Complexity is quadratic $\rightarrow$ cheap

1.2.6 Algorithm 6 (Extended Euclidean Algorithm)

input : $a, b \in \mathbb{N}$

output: $d = \gcd(a, b)\quad$ and $s, t \in \mathbb{Z}\quad$ such that $d = s \cdot a + t \cdot b$

(1) $r_0 := a, r_1 := b, s_0 := 1, t_0 := 0, s_1 := 0, t_1 = 1$

(2) for $i = 1, 2, ...\quad$ perform steps (3) - (5)

(3) if $r_i = 0\quad$ set $d = |r_{i-1}|$
$s := sgn(r_{i-1}) \cdot s_{i-1}$,
$t := sgn(r_{i-1}) \cdot t_{i-1}$

(4) division with remainder:
$r_{i+1} = r_{i-1} - q_i \cdot r_i,\quad$ with $|r_{i+1}| \leq \frac{1}{2}|r_i|$

(5) set $s_{i+1} := s_{i-1} - q_i \cdot s_i$,
$t_{i+1} := t_{i-1} - q_i \cdot t_i$

Justification : $r_i = s_i \cdot a + t_i \cdot b$ throughout
Application: $m, x \in \mathbb{N}$ such that m, x co-prime (i.e. $\gcd(x, m) = 1$)
Algorithm 6 yields: $1 = s \cdot x + t \cdot m \Rightarrow s \cdot x \equiv 1\ mod\ m$. So obtain inverse of $\bar{x} \in \mathbb{Z}/(m)$

1.3 Primality testing

Let $\mathbb{P} \subseteq \mathbb{N}$ be the set of prime numbers.
Challenge: Given $n \in \mathbb{N}$ decide if $n \in \mathbb{P}$
Naive Method: Trivial division by $m \leq \lfloor\sqrt{n}\rfloor$.
Running time is exponential in $l(n)$. Even when restricted to division by prime numbers, need approximatily $\frac{\sqrt{n}}{|n|\sqrt{n}}$ trivial divisions (prime number theorem)
$\rightarrow$ hardly any better!
Reminder: (arithmetic modulo m)
G finite group $\Rightarrow \forall a \in G \quad a^{|G|} = 1$ Fermat's little theorem
For $G = (\mathbb{Z}/(p)^{\times} \quad a^{p-1} \equiv 1 \quad (mod\, p) \in \mathbb{P} \quad \forall a \in \mathbb{Z}$ with $p \nmid a$
Infact $(\mathbb{Z}/(p))^{\times} \cong Z_{p-1}$ is cyclic
For $m = p_1^e, ...p_r^{e_r}$ with $p_i \in \mathbb{P},\ e_i \in \mathbb{N}_{>0}$:
$\mathbb{Z}_{(m)} \cong \mathbb{Z}_{(p_i^{e_i})} \oplus ... \oplus \mathbb{Z}_{(p_r^{e_r})} \Rightarrow \mathbb{Z}^x_{(m)} \cong \mathbb{Z}^x_{(p_i^{e_i})} \times ... \times \mathbb{Z}^x_{(p_r^{e_r})}$
what is $\mathbb{Z}_{(p^e)}$for $p \in \mathbb{P},\ e \in \mathbb{N}_{>0}$?

1.3.1 Theorem 1 (Cyclic group)

Let $p \in \mathbb{P}\ odd\ \ e \in \mathbb{N}_{>0} \Rightarrow (\mathbb{Z}_{(p^e)})^{\times} = Z_{(p-1)\cdot p^{e-1}}$ cyclic
Proof:
$(\mathbb{Z}_{(p^e)})^{\times} \cong Z_{p-1} \Rightarrow \exists z \in \mathbb{Z} : order(z + p\mathbb{Z}) = p - 1$
Set $a = \bar{z}^{p^{e-1}} \in (\mathbb{Z}_{(p^e)})^{\times} =: G$
$\quad a^{p-1} = \bar{z}^{(p-1)\cdot p^{e-1}} = \bar{z}^{|a|} = 1$
On the other hand, take $i \in \mathbb{Z}$ such that
$a^i = 1 \Rightarrow z^{i \cdot p^{e-1}} \equiv 1 \quad (mod\, p) \Rightarrow (p-1)|(i - p^{e-1}) \Rightarrow (p-1)|i.$
So $ord(a) = p - 1$.
Now consider $b = (p \bar{+} 1) \in G$
Claim: $ord(b) = p^{e-1}$
Proof by induction on $k \in N_{>0}$ that $(p+1)^{p^{k-1}} \equiv p^k + 1 \quad (mod\, p^{k+1})$
$k = 1 \quad \checkmark$
$k \rightarrow k+1$: By induction have $(p+1)^{p^{k-1}} = 1 + p^k + x \cdot p^{k+1}, \quad x \in \mathbb{Z}$
Compute: $(p+1)^{p^k} = ((1 + p^k) + x \cdot p^{k+1})^p = \sum\limits_{i=0}^{p} \binom{p}{i}(i + p^k)^{p-i} \cdot x^i \cdot p^{i\cdot(k+1)}$

$\underset{\text{Only 0-th summand}}{\equiv} (i + p^k) = \sum\limits_{i=0}^{p} \binom{p}{i} p^{i\cdot k} \underset{p\ odd}{\equiv} 1 + p^{k+1} \quad (mod\, p^{k+2}) \quad \checkmark$
For $k = e : (p+1)^{p^{e-1}} \equiv 1 \quad (mod\, p^e) \Rightarrow b^{p^e} = 1 \Rightarrow ord(b)|p^{e-1}$
But $(p+1)^{p^{e-2}} \equiv p^{e-1} + 1 \quad (mod\, p^e) \Rightarrow b^{p^{e-2}} \neq 1 \in G$
So $ord(b) = p^{e-1}$
Claim: $ord(a \cdot b) = (p-1)p^{e-1}$ $\quad (\Rightarrow$ Theorem)
Let $(a \cdot b)^i = 1 \in G$ with $i \in \mathbb{Z}$
Then $1 = (a \cdot b)^{i\cdot(p-1)} = (a^{p-1})^i \cdot b^{i\cdot(p-1)} = b^{i\cdot(p-1)} \Rightarrow p^{e-1}|i \cdot i(p-1) \Rightarrow p^{e-1}|i$
Also $1 = (a \cdot b)^{p^{e-1}\cdot i} = a^{p^{e-1}} \Rightarrow (p-1)|p^{e-1} \cdot i \Rightarrow (p-1)|i \rightarrow (p-1) \cdot p^{e-1}|i \quad \square$
Reminder: $(\mathbb{Z}/(2^e))^{\times} \cong Z_2 \times Z_2^{e-2} \quad (e \geq 2)$

1.3.2 Algorithm 2 (Fermat Test)

input : $n \in \mathbb{N}_{>0} odd$

output: "$n \notin \mathbb{P}$" or "probably $n \in \mathbb{P}$"

(1) Choose $a \in 2, ..., n-1$ randomly

(2) Compute $a^{n-1}\ mod\ n$

(3) If $a^{n-1} \not\equiv 1 \quad (mod\ n)$ then return "$n \notin \mathbb{P}$"
otherwise return "probably $n \in \mathbb{P}$"

Not very satisfying. Is this fast?

1.3.3 Algorithm 3 (fast exponentiation)

input : $a \in G \quad G$ is a monoid, $e \in \mathbb{N}, e = \sum\limits_{i_0}^{n-1} e_i 2^i, e_i \in \{0,1\}$

output: $a^e \in G$

(1) Set $b := a, \quad y := 1$

(2) For $i = 0, ..., n-1$ perform (3) - (4)

(3) if $e_i = 1$ set $y := y \cdot b$

(4) set $b := b^2$

(5) return y

this requires $O(l(e))$ operations in G
For $G = (\mathbb{Z}/(n)_i)$, each multiplication requires $O(l(n)^2)$ bit operations
$\Rightarrow$ Fermat test requires $O(l(n)^3)$ bit operations $\rightarrow$ cubic complexity $\rightarrow$ "fast"!
Example:
$n = 561 = 3 \cdot 11 \cdot 17 \quad$ For $a \in \mathbb{Z} \quad$ with $\gcd(a,n) \Rightarrow$ have $a^{n-1} = (a^2)^{280} \equiv 1\ mod\ 3$
$a^{n-1} \equiv 1 \quad (mod\ n)$ Fermat's test says "probably $n \in \mathbb{P}$" in 57% of cases.
$n = 2207 \cdot 6619 \cdot 15443$: output "probably $n \in \mathbb{P}$" in 99,93% of cases.

1.3.4 Definition 4 (pseudo-prime, witness, Carmichael numbers)

Let $n \in N_{>1} odd, \quad a \in 1, ..., n-1$

(a) n is pseudo-prime to base a if $a^{n-1} \equiv 1 \quad (mod\, n)$

(b) otherwise a is called a witness of composition of n

(c) If $n \notin \mathbb{P}$ but $a^{n-1} \equiv 1\; mod\; n \quad \forall a$ with $\gcd(n, a) = 1$
then n is called a Carmichael number.
There are ∞ Carmichael numbers

1.3.5 Proposition 5 (Number of witnesses)

Let $n \in N_{>1}, \quad odd \notin \mathbb{P} \quad$ not Carmichael
$\Rightarrow |\{a \in \mathbb{Z} \mid 0 < a < n$ a witness of composite of $n\}| > \frac{n-1}{2}$
Proof: Consider
$\phi : (\mathbb{Z}/(n))^{\times} =: G \to G, \quad \bar{a} \mapsto \bar{a}^{n-1}$
group homomorphism. By assumption,
$|im(\phi| > 1 \Rightarrow |Ker(\phi)| \leq \frac{|a|}{2} < \frac{n-1}{2}$
$\Rightarrow |\{a \in \mathbb{Z} | 0 < a < n$ a witness of composite of $n\}| > \frac{n-1}{2} \qquad \square$

Miller-Rabin Test

1.3.6 Proposition 6 (inference from Fermat)

Let $p \in \mathbb{P}\; odd, \quad a \in \{1, ..., (p-1)\} \quad$ write $p - 1 = 2^k \cdot m \quad$ with $m\; odd$ Then:
$a^m \equiv 1 \quad (mod\, p) \quad$ or
$\exists i \in \{0, ..., k-1\} : a^{2^i \cdot m} \equiv -1 \quad (mod\, p)$
Proof:
Little Fermat: $\bar{a}^{2^k \cdot m} = 1 \in \mathbb{F}_p$
Assume $\bar{a}^m \neq 1$ take i maximal such that:
$\bar{b} = \bar{a}^{2^i \cdot m} \neq 1 \Rightarrow \bar{b}^2 = 1 \Rightarrow \bar{b} \in \mathbb{F}_p$ is a zero of $x^2 - 1 \in \mathbb{F}_p[x] \Rightarrow \bar{b} = -1$

1.3.7 Algorithm 7 (Miller-Rabin-test)

input : $n \in \mathbb{N}_{>1}, odd$

output: either "$n \notin \mathbb{P}$" or "probably $n \in \mathbb{P}$" $\rightarrow$ Monte Carlo Algorithm.

(1) write $n - 1 = 2^k \cdot m$ with m odd

(2) Choose $a \in \{2, ..., n-1\}$ randomly

(3) Compute $b := a^m \; mod \; n$

(4) if $(b \equiv \pm 1 \pmod n)$
return "probably $n \in \mathbb{P}$"

(5) for $(i = 0, ..., k-1)$ do steps (6) - (7)

(6) set $b := b^2 \pmod n$

(7) if $(b \equiv -1 \pmod n)$
return "probably $n \in \mathbb{P}$"

(8) return $n \notin \mathbb{P}$"

1.3.8 Definition 8 (strong pseudo-prime / witness)

Let $n \in \mathbb{N}_{>1}, \; odd \quad a \in \{1, ..., n-1\}$

(a) n is called a strongly pseudo-prime to base a if Proposition 6 holds for a and p replaced by n.

(b) Otherwise a is called a strong witness of composition of n.

Example
Let $n \in \mathbb{N}_{>1}, \quad \mathbb{P} \; odd$
$a = 2$ strong witness if $n < 2047$ (including 561)
2 or 3 strong witness if $n < 1373653$
2,3 or 5 strong witness if $n < 25326001$

1.3.9 Theorem 9 (Bit-complexity of Algorithm 7)

(a) Algorithm 7 requires $O(l(n)^3)$ bit operations. $\rightarrow$ "qubic complecity" $\rightarrow$ fast!

(b) if $b \in \mathbb{P}$ then Algorithm 7 returns "probably $b \in \mathbb{P}$" $\rightarrow$ no false positives.

(c) if $n \notin \mathbb{P}$ then more than half of the numbers in $\{1, ..., n-1\}$ are strong witnesses.

Proof:

(a) Step 1 takes $O(l(n))$ bit operations:
Using Algorithm 3, we need $O(l(n-1))$ multiplications in $\mathbb{Z}/(n)$ each requiring $O(l(n)^2)$ bit operations.

(b) Proposition 6

(c) split in three cases:

Case 1: n is not a Carmichael number. $\xRightarrow[\text{Prop 5}]{}$ more than half of all numbers are.
Fermat witness thus also strong witness.
Case 2: $n = p^r \cdot l$ with $p \in \mathbb{P} \quad r > 1 \quad l \in \mathbb{N}_{>0} p \nmid l$
Theorem 1 $\exists x \in Z$ such that $x^p \equiv 1 \; mod \; p^r \quad x \neq 1 \; mod \; p^r$
Chinese remainder theorem: $\exists a \in \mathbb{Z}$ such that $a \equiv x \; mod \; p^r \quad a \equiv 1 \; mod \; l$
So $\bar{a}^p = 1 \in \mathbb{Z}_(n) \Rightarrow \bar{a}^n = 1 \Rightarrow \bar{a} \in (\mathbb{Z}/(n))^\times$
i.e. $\gcd(n, a) = 1$ if $\bar{a}^{n-1} = 1$ then $\bar{a} = 1$
But $a \equiv x \neq 1 \; mod \; p^r$ so $\bar{a}^{n-1} \neq 1$ hence n is not Carmichael $\rightarrow$ Case 1.
Case 3: n is a Carmichael number. By Case 2 have $n = p \cdot l$ with $p \in \mathbb{P} \quad p \nmid l \quad l \geq 3$
n Carmichael: $\forall a \in \mathbb{Z}$ with $\gcd(a, n,) = 1$
have $a^{2^k \cdot m} \equiv 1 \; mod \; n$ (where $n - 1 = 2^k \cdot m$)
$a^{2^k \cdot m} \equiv 1 \; mod \; p$ Take j minimal such that
$a^{2^j \cdot m} \equiv 1 \; mod \; p \forall a \in \mathbb{Z}$ such that $\gcd(a, n) = 1$
so $0 \leq j \leq l$ in fact, $j > 0$ since $(-1)^{2^0 \cdot m} = -1$ with m *odd.*
Consider the subgroup $H := \{\bar{a} \in \mathbb{Z}/(n) | \bar{a}^{2^{j-1} \cdot m} \in \{1, -1\} \subseteq (\mathbb{Z}/(n))^\times\}$
Let $a \in \{1, ..., n-1\} \quad \gcd(n, a) = 1 \quad a$ not a strong witness.
Claim 1: $\bar{a} \in H$
das da **Case 3.1:** $\bar{a}^{2^{j-1} \cdot m} = 1 \Rightarrow \bar{a} \in H$
Case 3.2: $a^{2^{j-1} \cdot m} \neq 1 \; mod \; n \qquad a^m \neq 1 \; mod \; n$
$\xRightarrow[a \text{ nonwitness}]{} \exists i$ such that $\underbrace{a^{2^i \cdot m} \equiv -1 \; mod \; n}_{*}$
$\Rightarrow a^{2^i \cdot m} \equiv -1 \; mod \; p \xRightarrow[\text{def of j}]{} i < j$
if $i < j - 1$ then $a^{2^{j-1} \cdot m} = (a^{2^i \cdot m})^{2^{j-1-i}} \equiv (-1)^{2^{j-1-i}} = 1 \; mod \, n$
$\xRightarrow[\text{with } *]{}$ not in case 3.2
Claim 2: $H \subseteq (\mathbb{Z}_(n))^\times$ proper subgroup.
By definition of $j \exists x \in \mathbb{Z}$ such that $x^{2^{j-1} \cdot m} \neq 1 \; mod \; p$
Chinese remainder: $\exists a \in \mathbb{Z}$ such that
$a \equiv x \; mod \; p \qquad a \equiv 1 \; mod \; l \Rightarrow a^{2^{j-1} \cdot m} \neq 1 \; mod \; p \equiv 1 \; mod \; l \Rightarrow \bar{a} \notin H$
Claim 2 ✓
It follows that $|H| \leq \frac{|(\mathbb{Z}/(n))^\times|}{2} < \frac{n-1}{2}$
so the number of witnesses is $\geq n - 1 - |H| > \frac{n-1}{2}$ $\square$

Remarks:

(a) A more careful analysis shows that $2\frac{3}{4}$ of all candidates are strong witnesses

(b) Calling Algorithm 7 repeatedly decreases the probability of false positives. Running time for prescribed error probability p is $O(\lg(p^{-1} \cdot l(n)^3))$
(Independence assumptions!)

Connection with Riemann hypothesis
Let $n \in \mathbb{N}_{>0}$ $\quad \bar{X} : (\mathbb{Z}/(n))^\times \to \mathbb{C}^x$ group homomorphism
$X : \mathbb{Z} \to \mathbb{C}, a \mapsto \begin{cases} \bar{X}(\bar{a}) & \text{if } \gcd(a,n) = 1 \quad \text{for } (\bar{a} = a + n\mathbb{Z}) \\ 0 & \text{otherwise} \end{cases}$
"residence class character *mod* n
$Ex : n = 1 \Rightarrow X(a) = 1 \forall a \in \mathbb{Z}$
Divichlet L-series:
$L_X(s) = \sum\limits_{n=1}^{\infty} \frac{X(n)}{n^s}$ converges for $s \in \mathbb{C}$ until $Re(s) > 1$
$L_X(s)$ extends to a meromorphic function on $\mathbb{C} \mapsto$ "Divichlet L-function".
For $n = 1 : L_X(s) = \zeta(s)$ Riemann Zeta-function.
Euler Product:
From $(1 - X(p) \cdot p^{-s})^{-1} = \sum\limits_{i=0}^{\infty} (X(p) \cdot p^{-s})^i = \sum\limits_{i=0}^{\infty} \frac{X(p^i)}{p^{is}}$ $\quad$ derive $L_X(s) = \prod\limits_{p \in \mathbb{P}} \frac{1}{1 - X(p) \cdot^{-} s}$
Generalized Riemann hypothesis (GRH):
For X residue class character, $s \in \mathbb{C}$
with $L_X(s) = 0, \quad 0 < Re(s) < 1$ $\qquad$ ("critical strip")
then $Re(s) = \frac{1}{2}$
For $X = 1 \to$ ordinary Riemann hypothesis.

1.3.10 Theorem (Arkeny & Bach)

GRH $\Rightarrow \forall X \neq 1$ residence class character
$\exists p \in \mathbb{P} : X(p) \neq 1, p < 2 \ln(n)^2$

Let $H \subsetneqq (\mathbb{Z}/(n))^\times =: G$ proper subgroup.
Choose $N \subsetneqq G$ maximal proper subgroup such that $H \subseteq N \Rightarrow G/N$ cyclic.
$\bar{X} : G \mapsto \mathbb{C}^x$ with $N = Ker(\bar{X}) \Rightarrow H \subseteq Ker(\mathbb{X})$
$\xRightarrow[\text{GRH, Thm1}]{} \exists p \in \mathbb{P} : p + n\mathbb{Z} \notin H, p < 2 \cdot \ln(n)^2$
Corollary: Assume GRH.
Let $n \in \mathbb{N}_{>1}$ $\quad \mathbb{P}$ *odd* Then there is a strong witness a of compositeness of n with $a < 2 \cdot \ln(n)^2$.
$\to$ Obtain deterministic primality test with time $O(\ln(n)^5)$ bit operations.

AKS-test
A deterministic polynomial time primality test $\to$ "holy grail"
Agrawal, Kayal, Saxena: PRIMES is in P, Annals of Mathematics, 2004.

1.3.11 Proposition 10 (modulo over ideals)

Let $n \in \mathbb{P} \quad a \in \mathbb{Z} \Rightarrow (x+a)^n \equiv x^n + a \quad (mod\ n)$
where x is a indeterminate and for $r \in \mathbb{N}$:

$$(x+a)^n \equiv (x^n + a) \quad (mod\ (n, x^r - 1)) \tag{1}$$

(i.e. $(x+a)^n - (x^n + a) = n \cdot f + (x^r - 1) \cdot g \quad$ with $f, g \in \mathbb{Z}[x]$)
Proof:
$(x+a)^n = \sum_{i=0}^{n} \binom{n}{i} \cdot x^i a^{n-1} \qquad$ (where $\binom{n}{i}$ is a multiple of n for $0 < in$)
$\equiv x^n + a^n \qquad (\leftarrow$ little Fermat)
$\equiv x^n + a \qquad$ (1) follows by weakening this.
Cost analysis for checking (1) with l = length(n).
Using Algorithm 3, need $O(l)$ multiplications in $\mathbb{Z}[x]/(n, x^r - 1) =: R$
Elements of R are represented as polynomials of degree $< r$,
coefficients between 0 and n.
Multiply polynomials: $O(r^2)$ operation in $\mathbb{Z}/(n) : O(r^2 \cdot l^2)$
since $x^{r+k} \equiv x^k\ mod\ x^r - 1$,
add coefficients of x^{r+k} of product polynomial to coefficients $x^k : O(r \cdot l)$
Total for checking (1): $O(r^2 \cdot l^3)$ bit operations.
Reduction $mod\ x^r - 1$ is just for keeping the cost under control.
The following is part of AKS-test:

1.3.12 Algorithm 11 (Test for perfect power)

input : $n \in \mathbb{N}_{>1}$

output: $m, e \in \mathbb{N} \quad e > 1$ such that $n = m^e \quad$ or "n is not a perfect power"

(1) for $(e = 2, ..., \lfloor \lg(n) \rfloor)$ perform (2) - (7) //possible exponents

(2) set $m_1 = 2, m_2 = n$ //initialize interval $[m_1, m_2]$ for searching $\sqrt[e]{n}$

(3) while($m_1 \leq m_2$) do (4) - (7)

(4) set $m = \lfloor \frac{m_1 + m_2}{2} \rfloor$ // bisect interval

(5) if $m^e = n$ return m, e

(6) if $m^e > n$ set $m_2 = m - 1$

(7) if $m^e < n$ set $m_1 = m + 1$

(8) return "not a perfect power"

Cost: (for l = length(n))
Compute $m^e : O(\lg(l) \cdot l^2)$ (abort computation once the result exceeds n)
Number of passages through inner loops $\leq \lg(n)$
Number of passages through outer loops $\leq \lg(n)$
Total cost of Algorithm 11: $O(l^4 \cdot \lg(l))$

1.3.13 Algorithm 12 (AKS-test)

input : $n \in \mathbb{N}_{>1}$ of length $l = \text{length}(n,) = \lfloor \lg(n) \rfloor + 1$

output: "$n \in \mathbb{P}$" or "$n \notin \mathbb{P}$"

(1) if (n is a perfect power)
return "$n \notin \mathbb{P}$"

(2) find $r \in \mathbb{N}_{>1}$ minimal such that $\quad r|n \quad \vee \quad n^i \not\equiv 1 \ (mod\ r) \qquad \forall i = 1, ..., l^2$
//exhaustive search (we will show that $r \leq l^5$)

(3) if $r|n$
if $(r = n)$ return "$n \in \mathbb{P}$"
if $(r < n)$ return "$n \notin \mathbb{P}$"

(4) for $a = 1, 2..., \lfloor \sqrt{r} \cdot l \rfloor$ do (5)

(5) if $((x+a)^n \not\equiv (x^n + a) \qquad (mod\ (n, x^r - 1)))$
return "$n \notin \mathbb{P}$"

(6) return "$n \in \mathbb{P}$"

1.3.14 Lemma 13 (Least common multiple)

For $n \in \mathbb{N}_{>0}$ have $\lambda(n) := lcm(1, 2, ...n) \geq 2^{n-2}$

Proof: For $f = \sum\limits_{i=0}^{m} a \cdot x^i \in \mathbb{Z}(x) \quad a_i \in \mathbb{Z}$

$\Rightarrow \int\limits_0^1 f(x)dx = \sum\limits_{i=0}^{m} \frac{a_i}{i+1} = \frac{k}{\lambda(m+1)}$

with $k \in \mathbb{Z}$. Consider $f_m = x^m \cdot (1-x)^m$

For $0 < xy1$:

$0 < f_m(x) \leq 4^{-m}$

$\Rightarrow 0 < \int\limits_0^1 \underbrace{f_m(x)}_{\frac{k_m}{\lambda(2m+1)}} dx \leq 4^{-1}$

$\lambda(2 \cdot m + 1) \geq k_m \cdot 4^m \geq 4^m$

For $n \in \mathbb{N}_{>0} \lambda(n) \geq \lambda(2 \cdot \lfloor \frac{n-1}{2} \rfloor + 1) \geq 4^{\lfloor \frac{n-1}{2} \rfloor} \geq 4^{\frac{n-1}{2}} = 2^{n-2} \qquad \square$

Corollary: (not related to AKS)

For $n \in \mathbb{M}$

$\pi(n) := |\{p \in \mathbb{P} | p \leq n\}| \geq \frac{n-2}{\lg(n)}$

Proof:

$2^{n-2} \leq \lambda(n) = \prod\limits_{p \in \mathbb{P}, p \leq n} p^{\lfloor \log_p(n) \rfloor} \leq \prod\limits_{p \leq n} p^{\log_p(n)} = n^{\pi(n)} = 2^{\lg(n)\pi(n)} \qquad \square$

Prime number theorem:
$\lim\limits_{n\to\infty} \frac{\pi(n)}{n/\ln(n)} = 1$
Interpretation:
The average distance of two primes around some value $x \in \mathbb{R}_{>1}$ is $\ln(x)$

1.3.15 Lemma 14 (Property of r in Algorithm 12)

In Algorithm 12, have $r \leq l^5$
Proof:
if $r < l^5 \Rightarrow \forall k \in k \in \{2, ..., l^5\} : \exists i \in \{1, ..., l^2\}$
$n^i \equiv 1 \pmod{k}$
$\Rightarrow k | \prod\limits_{i=1}^{l^2} (n^i - 1)$
$\Rightarrow \lambda(l^5) | \prod\limits_{i=1}^{l^2} (n^i - 1)$
$\xRightarrow[Lemma 13]{} 2^{l^5-2} < \prod\limits_{i=1}^{l^2} n^i = n^{\frac{l^2(l^2+1)}{2}}$
$\Rightarrow l^5 - l^3 < 4$ not true since $l \geq 2$ $\square$

1.3.16 Theorem 15 (Bit-Complexity of Algorithm 12)

Algorithm 12 requires $O(l^{16.5})$ bit operations (”polynomial complexity”)
Proof:

Step(1): $O(l^4 \cdot \lg(l))$✓

Step(2): For each r need:

- test $r|n : O(l^2)$
- compute all $n^i \; mod \; r : O(l^2 \cdot \lg(r)^2) \underset{Lemma 14}{\leq} O(l^2 \cdot \lg(l)^2)$

Step(3): $O(1)$

Step(4): $O(\sqrt{r} \cdot l \cdot r^2 \cdot l^3) \underset{Lemma 14}{\leq} O(l^{16,5})$ $\square$

Reminder: There is a variant of Algorithm 12 with running time $\tilde{O}(l^6)$,i.e., $O(l^6 \cdot \lg(l)^m)$ with $m \in \mathbb{N}$.
Correctness:
For $r \in \mathbb{N}_{>0}$ and $p \in \mathbb{P}$ write $I(r,p) := \{m, f) \in \mathbb{N} \times \mathbb{F}_p[x] | f(x)^m \equiv f(x^m) \; mod \; (x^r - 1)\}$
”m is introspective for f and r”.
Example: Proposition 10 says that:

$$(p, x + \bar{a}) \in I(r,p) \text{for } a \in \mathbb{Z} \quad r \in \mathbb{N}_{>0} \quad p \in \mathbb{P} \tag{1}$$

1.3.17 Lemma 16 (Rules)

(a) $(m, f), (m', f) \in I(r, p) \Rightarrow (m \cdot m', f) \in I(r, p)$

(b) $(m, f), (m, g) \in I(r, p) \Rightarrow (m, f \cdot g) \in I(r, p)$

(c) $(m \cdot p, f) \in I(r, p), p \nmid r \Rightarrow (m, f) \in I(r, p)$

Proof:

(a) $f(x)^{m \cdot m'} \equiv f(x^m)^{m'} \; mod \; (x^r - 1)$
$f(x^m)^{m'} \equiv f(x^{m \cdot m'}) \; mod \; (x^{m \cdot r} - 1)$
But $(x^r - 1 | (x^{m \cdot r} - 1))$

(b) $(f \cdot g)(x)^m = f(x)^m \cdot g(x)^m \equiv f(x^m) \cdot g(x^m) = (f \cdot g) \cdot (x^m) \; mod \; (x^r - 1)$

(c) $(f(x)^m)^p \equiv f((x^m)^p) \underset{Frobeniushomomorphism}{=} (f(x^m))^p \; mod \; (x^r - 1)$
$\Rightarrow (x^r - 1)|((f(x)^m)^p - f(x^m)^p) \underset{Frobeniushomomorphism}{=} (f(x)^m - f(x^m))^p$
$p \nmid r \Rightarrow x^r - 1$ is square free. So
$(x^r - 1)|(f(x)^m) - f(x^m)) \Rightarrow (m, f) \in I(r, p)$ $\square$

1.3.18 Theorem 17 (Correctness of Algorithm 12)

Algorithm 12 is correct.
Proof:
If the algorithm terminates in step(1),(3) or (5), it is correct. To show: If it terminates in step(6) it is correct, i. e. $n \in \mathbb{P}$
Claim 1: $\exists p \in \mathbb{P} : p|n \quad p \neq 1 \; mod \; r \quad p > r$
Indeed if all prime divisors of n were $\equiv 1 \; mod \; r$ then $n \equiv 1 \; mod \; r$ contradiction to step(2). All prime divisors of n are $> r$ by step (2) and (3) ✓
Steps(2) and (3) imply that $\gcd(n, r) = 1 \Rightarrow G := < \bar{n}, \underbrace{\bar{p}}_{p \; mod \; r} > \subseteq (\mathbb{Z}/(r))^\times$

Step(2): $ord(\bar{n}) > l^2 \Rightarrow l^2 < |G| < r$ (2)

Set $s := ord(\bar{p} \in G) \Rightarrow r|(p^s - 1)$ with $q := p^s \Rightarrow r||\mathbb{F}_q^\times| \Rightarrow \exists \zeta \in \mathbb{F}_q$ r-th root of 1
Set $k := \lfloor \sqrt{r} \cdot l \rfloor \quad m := (\frac{n}{p})$
By (1) $(p, x + \bar{a}) \in I(r, p)$ with $\bar{a} \in \mathbb{F}_p$
By step(4), have $(n, x + \bar{a}) \in I(r, p)$
For $\underline{e} = e_0, ..., e_k \in \mathbb{N}_0$ set $f_{\underline{e}} := \prod_{a=0}^{k} (x + \bar{a})^{e_a}$
Lemma 16 (b): $(p, f_{\underline{e}}) \in I(r, p)$
$(n, f_{\underline{e}}) \in I(r, p)$
$\xRightarrow[Lemma16(c)]{} (m, f_{\underline{e}}) \in I(r, p)$
$\xRightarrow[Lemma16(a)]{} \forall s, t \in \mathbb{N}_0 : (p^s \cdot m^t, f_{\underline{e}}) \in I(r, p)$

$\Rightarrow f_{\underline{e}}(\zeta^{p^s \cdot m^t}) = f_{\underline{e}}(\zeta)^{p^s \cdot m^t}$ (3)

Set $H :=< \zeta + \bar{a} | a \in \{0, ..., k\} > \subseteq \mathbb{F}_q^\times$
($\zeta \notin \mathbb{F}_p$ since $r \nmid (p-1)$ by Claim 1)
Consider: $T := \{(e_0, ..., e_k) \in \mathbb{N}_0^{k+1} | \sum\limits_{a=0}^{k} e_a < |G|\}$
$\Phi : T \mapsto H, (e_0, ..., e_k) \mapsto f_{\underline{e}}(\zeta) = \prod\limits_a (\zeta + \bar{a})^{e_a} \in H$
Claim 2: Φ is injective.
Indeed, take $(\underline{e}), (\underline{\hat{e}}) \in T$ such that $\Phi(\underline{e} = \Phi(\underline{\hat{e}}))$
$\Rightarrow \forall s, t \in \mathbb{N}_0 : f_{\underline{e}}(\zeta^{p^s \cdot m^t}) \underset{(3)}{=} f_{\underline{e}}(\zeta)^{p^s \cdot m^t} = f_{\underline{\hat{e}}}(\zeta)^{p^s \cdot m^t} \underset{(3)}{=} f_{\underline{\hat{e}}}(\zeta^{p^s \cdot m^t})$
$f_{\underline{e}} - f_{\underline{\hat{e}}}$ has roots ζ^e with $e \in G$ since $G =< \bar{p}, \bar{m} >$
These are all distinct (since ζ is primitive)
But $deg(f_{\underline{e}} - f_{\underline{\hat{e}}}) < |G|$ So $f_{\underline{e}} - f_{\underline{\hat{e}}} = 0$
Since $k \leq \sqrt{r} \cdot l \underset{(2)}{<} r < p$ the $(x + \bar{a})$ with $a \in \{0...k\}$ are primitive distinct.
So $(\underline{e}) = (\underline{\hat{e}})$ ✓
So is $|H| \geq |T|$?
Let M be the set of all $\{x_0, ..., x_k\} \subseteq \{1, ..., |G| + k\}$
with $x_0 < x_1 < ... < x_k$
For $\{x_0, ..., x_k\} \in M$ define $(e_0, ..., e_k) \in \mathbb{N}_0^{k+1}$ by $e_a = x_a - x_{a-1}$ with $x_{-1} := 0$
$\Rightarrow \sum\limits_{a=0}^{k} e_a = \sum\limits_{a=0}^{k} (x_a - x_{a-1} - 1) = x_k - (k+1) < |G|$
Obtain injection $M \Leftrightarrow T$
So $|H| \geq |T| \geq |M| = \binom{|G|+k}{k+1} \underset{(2)}{\geq} \binom{\lfloor l\sqrt{|a|} \rfloor + 1 + k}{k+1} = \binom{\lfloor l\sqrt{|a|} \rfloor + 1 + k}{\lfloor l\sqrt{|a|} \rfloor} \underset{(2)}{\geq} \binom{2 \cdot \lfloor l\sqrt{|a|} \rfloor + 1}{\lfloor l\sqrt{|a|} \rfloor}$

1.3.19 Lemma 18 (Property of binomial coefficients)

$\forall n \in \mathbb{N}_{>1} : \binom{2 \cdot n + 1}{n} > 2^{n+1}$
Proof:
$n = 2 :$
$\binom{5}{2} = 10 > 2^3$
$n - 1 \to n :$
$\binom{2 \cdot n + 1}{n} = \binom{2 \cdot n}{n-1} + \binom{2 \cdot n}{n} = \binom{2 \cdot n - 1}{n-2} + \binom{2 \cdot n - 1}{n-1} + \binom{2 \cdot n - 1}{n-1} + \binom{2 \cdot n - 1}{n} \geq 2 \cdot \binom{2 \cdot n - 1}{n-1} \underset{ind.}{>} 2 \cdot 2^n = 2^{n+1}$

Continuation of Proof of Theorem 17

$$|H| > 2^{\lfloor l \cdot \sqrt{|a|} \rfloor + 1} \geq 2^{l \cdot \sqrt{|a|}} \geq 2^{\lg(n) \cdot \sqrt{|a|}} = n^{\sqrt{|a|}} \tag{4}$$

Assume $n \notin \mathbb{P}$ By step (1) m is not a perfect power
$\Rightarrow$ the map $\mathbb{N}_0 \times \mathbb{N}_0 \mapsto \mathbb{N}$ $(s, t) \mapsto p^s m^t$ is injective.
Set $A := \{p^s m^t | s, t \in \{0, .., \lfloor \sqrt{a} \rfloor\}\} \subseteq \mathbb{N}$
$\Rightarrow |A| = (\lfloor \sqrt{|a|} \rfloor + 1)^2 > |G|$
Since $G =< \bar{p}, \bar{m} > \subseteq (\mathbb{Z}/(r))^\times$ this implies that $\exists n, \hat{n} \in A$
such that $n \neq \hat{n}$ but $b \equiv \hat{n}\ (mod\ r)$.
Let $h \in H \Rightarrow h = f_{\underline{e}}(\zeta)$ with $(\underline{e}) \in \mathbb{N}_0^{k+1} \Rightarrow h^n \underset{(3)}{=} f_{\underline{e}}(\zeta^n) \underset{n \equiv \hat{n}\ (mod\ r)}{=} f_{\underline{e}}(\zeta^{\hat{n}}) \underset{(3)}{=} h^{\hat{n}}$

So the polynomial $Y^n - Y^{\hat{n}} \in \mathbb{F}_q[Y]$ has all elements of H as zeros.
But $deg(Y^n - Y^{\hat{n}}) \leq max\{n, \hat{n}\} \leq (p \cdot m)^{\lfloor\sqrt{|G|}\rfloor} \leq n^{\sqrt{|G|}} < |H|$
$\Rightarrow$ contradiction since $Y^n - Y^{\hat{n}} \neq 0$ $\square$

1.4 Cryptology

A ("Alice") wants to send a message to B ("Bob") such that an eavesdropper E ("Eve") can not read the clear message. So A and B encrypt the message.

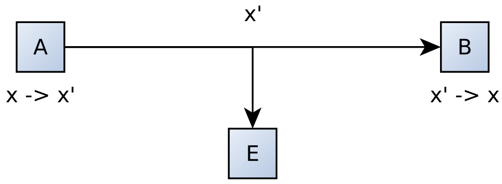

Figure 2: Scheme of eavesdropping

Symmetric-key cryptography
A and B share secret keys for encryption ($x \mapsto x'$) and decryption ($x' \mapsto x$) Only A and B know the keys.
Example: AES approved by the US government in 2002
Application:

- sending messages
- encrypt files (A=B)

Problem: Key exchange between A and B

Public-key cryptography
Encryption-map $\phi : x \mapsto x'$ is made public by B, but decryption $\phi : x' \mapsto x$ is kept secret.
Advantage: No confidential key exchange.
Disadvantages:

- more costly than symmetric key cryptography
- doubt weather E can reconstruct ϕ^{-1} from ϕ with enough computing power

Applications:

- sending messages
- exchange of symmetric keys
- authentication: Together with x, B sends $\phi^{-1}(x)$
  (or $\phi^{-1}|$ Part of x together with date). A verifies by applying ϕ.
  Better: challenge-response-protocol.

Examples: RSA, elliptic curve

RSA

(1) B chooses $p, q \in \mathbb{P}$ large ($> 100 digits$)
with $p \neq q \quad n := p \cdot q$

(2) B chooses $e, f \in \mathbb{N}$ large such that $e \cdot f \equiv 1\ (mod\ \phi(n))$
with $\phi(n) = (p-1)(q-1)$

(3) B makes n, e public, keep f secret

(4) The message is encoded as an element $x \in \mathbb{Z}/(n)$

(5) A computes $\phi(x) = x^e = y \in \mathbb{Z}/(n)$ and sends y

(6) B receives y and computes $y^f = x \in \mathbb{Z}/(n)$

Comments on steps of RSA:
(6) Have $e \cdot f = a \cdot (p-1) \cdot (q-1) + 1 \quad$ with $a \in N_{>0}$
$y^f = x^{e \cdot f}$

Case 1: $q \nmid f, q \nmid x \Rightarrow x^{a(p-1)(q-1)} = (x^{\phi(n)})^a \underset{LittleFermat}{\equiv} 1^a = 1 \Rightarrow x^{e \cdot f} = x \qquad \checkmark$

Case 2: $p|x, q \nmid x \Rightarrow x^{e \cdot f} \equiv 0 \equiv x \quad (mod\ p)$
$x^{e \cdot f} \equiv x \quad (mod\ q)$ as above.

Case 3: $q|x$ As Case 2

$\Rightarrow$ Correctness of decryption
Cost:

(1) Finding p, q of length approximately l. Prime-number theorem: Gap between two primes of length $\approx l$ is $O(l)$
Using Miller Rabin with error probability 2^m. Expected cost of (1) is $O(m \cdot l^4)$ bit operations.

(2) Choose e co-prime to $\phi(n)$ obtain f = inverse $mod\ \phi(n)$ by extended euclidean Algorithm: $O(l^2)$

(5)(6) Fast exponentiation: $O(l^3)$

Security of RSA: p and q must be so large that factorization of a is "impossible". Assumption that factorization is expensive could not be shown! But could f be obtained without knowing p and q? The following algorithm gives a negative answer. It shows that the problem of breaking RSA is always basically factorization.
Remember: $\phi(n)|(e \cdot f - 1) =: m \leq n^2$

1.4.1 Algorithm 1 (finding a divisor)

Input : $n \in \mathbb{N}_{>2}$ $odd\ squarefree\ e \notin \mathbb{P}$ and $m \in \mathbb{N}_{>0}$ such that $\phi(n)|m \quad m \leq n^2$

Output: $d \in \mathbb{N}$ with $d|n \quad 1 < d < n$

(1) Choose $a \in \{2, ..., (n-2)\}$ randomly
set $k := m$

(2) If $d := \gcd(a, n) \neg 1$
return d

(3) Repeat steps (4) - (8) //while(true)

(4) compute $d := \gcd(n, a^k - 1)$

(5) If $d = 1$ go to (1)

(6) If $d < n$ return d

(7) if k is *odd* go to (1)

(8) set $k := \frac{k}{2}$

Correctness is clear. What about termination and running time?

1.4.2 Proposition 2 (Complexity of Algorithm 1)

Algorithm 1 terminates in expected time $O(l(n)^4)$ bit operations (Las Vegas Algorithm).
Proof:
Set $l :=$ length(n)
Have $n = \prod_{0=1}^{r} p_i$ with $p_i \in \mathbb{P}$ distinct.
$\phi(n) = \prod_{i=1}^{r} (p_i - 1) \mid m$ So initially all $(p_i - 1)$ divide k.
At some iteration it happens for the first time that $(p_i - 1) \nmid k$
Then $k \equiv \frac{p_1 - 1}{2} \quad (mod\ (p_1 - 1)) \Rightarrow a^k \equiv \pm 1 \quad (mod\ p_i)$ -1 occurs fore some a
For those j with $(p_j - 1)|k$ have $n^k \equiv 1 \quad (mod\ p_j)$
Consider the group homomorphism: $\phi_i (\mathbb{Z}/(n))^\times \mapsto (\mathbb{Z}/(p_1))^\times \times ... \times (\mathbb{Z}/(p_r))^\times$
$\bar{a} \mapsto (a^k\ mod\ p_1, ..., a^k\ mod\ p_r)$
The image of ϕ is a product of groups $\{\pm\}$ or $\{1\}$ depending whether $(p_i - 1) \nmid k$ or $(p_i - 1)|k$
Conclusion:
For at least half of all a's, $\phi(\bar{a})$ is neither (1,...,1) nor (-1,...,-1)
If $a^k \equiv \ \ 1 \quad (mod\ p_j)$ then $p_j \mid (a^k - 1) \Rightarrow p_j \mid d$
If $a^k \equiv -1 \quad (mod\ p_j)$ then $p_j \nmid (a^k - 1) \Rightarrow p_j \nmid d$
So for these a the algorithm is successful.
This means that the expected number of a's that need to be tested is ≤ 2

(Since $\sum_{i=1}^{\infty} i \cdot \left(\frac{1}{2}\right)^i = 2$ More generally for $0 < p < 1 : p \cdot \sum_{i=1}^{\infty} i \cdot (1-p)^{i-1} = \frac{1}{p}$)
Analysis of running time (in bit operations) for each a (using gcd is quadratic) leads to the claim. $\square$

Problems of RSA:

- How difficult is factorization of integers (lower bound?)
- decryption of some or all messages without having f?

2 Notes

- $\mathbb{N} := \mathbb{N}_0$
- $\lg(x) := \log_2(x)$
- $a \mid b$ a is divisible by b
  $a \nmid b$ a is not divisible by b
- $ord(a)$ order of a group element
  $n > 0$ minimal such that $a^n = e$ with neutral element e
  if no such n can be found, $ord(a) = \infty$
- Lagrange's theorem: every element in a finite group has finite order
- $char(A)$ the smallest positive n such that
  $\underbrace{1 + ... + 1}_{n \ \ summands} = 0$ with 1 as the multiplicative identity element
- $\mathbb{Z}/(m)$ Ring modulo m
  polynomial rings measure for "$<$" relations not abs value but max power
- Average number of bit operations for an increment:
  one operation for the last bit + 50% chance for one on the next bit + 25% on the following etc. $\Rightarrow$ Geometrical row
  $\Rightarrow$ on average two bit operations
- "Monte Carlo Algorithm": Always terminates in reasonable time but might yield false result.
- "Las Vegas Algorithm": If it terminates the result is correct. No deterministic running time.
- $lcm(a_1, ..., a_n)$ "least common multiple"
- $\underline{e}$ = vector of e's
- $\phi(n) := |(\mathbb{Z}/(n))^x|$ Euler totient function

2.1 Algebraic structures

• Group	$(G, *)$	
- one inner operation $(*)$:	$G \times G \mapsto G$	
- associativity:	$(a * b) * c = a * (b * c)$	$\forall a, b, c \in G$
- neutral element $e \in G$:	$a * e = e * a = a$	$\forall a \in G$
- inverse element $a^{-1} \in G$:	$a * a^{-1} = a^{-1} * a = e$	$\forall a \in G$
• Abelian group	$(G, *)$	
- $(G, *)$ is a group		
- commutativity:	$a * b = b * a$	$\forall a, b \in G$
• Finite group	$(G, *)$	
- associativity:	$(a * b) * c = a * (b * c)$	
- unambiguity of reduction:	$(a * x = a * x') \wedge (x * a = x' * a) \Rightarrow x = x'$	

$\Rightarrow x \mapsto x * a$ and $x \mapsto a * x$ is bijective
$\Rightarrow \exists x : a * x = a \Rightarrow$ neutral element
$\exists x : a * x = x \Rightarrow$ inverse element

- Cyclic group $(G, *)$
  - G is a group
  - G is generated by one Element: $G =< g >= \{g^n | n \in \mathbb{Z}\}$
  - not necessarily finite.
- Semi group $(S, *)$
  - one inner operation $(*)$: $S \times S \mapsto S$
  - associativity: $(a * b) * c = a * (b * c) \quad \forall a, b, c \in S$
- Field $(K, +, \cdot)$
  - two inner operations $(+, \cdot)$ such that:
    - $(K, +)$ is an abelian group with neutral element 0
    - $(K \backslash (0), \cdot)$ is an abelian group with neutral element 1
  - distributivity: $a \cdot (b + c) = a \cdot b + a \cdot c$
    $(a + b) \cdot c = a \cdot c + b \cdot c \quad \forall a, b, c \in K$
- Ring $(R, +, \cdot)$
  - $(R, +)$ is an abelian group
  - $(R, \cdot)$ is a semi group
  - distributivity: $a \cdot (b + c) = a \cdot b + a \cdot c$
    $(a + b) \cdot c = a \cdot c + b \cdot c \quad \forall a, b, c \in R$
- Commutative ring $(R, +, \cdot)$
  - $(R, +, \cdot)$ is a ring
  - commutativity for $(\cdot)$ $a \cdot b = b \cdot a \quad \forall a, b \in R$
- Unitary ring (ring with 1) $(R, +, \cdot)$
  - $(R, \cdot)$ is a semi group
  - $(R, \cdot)$ has a neutral element "1"
- Euclidean ring R
  - $\exists F : R \mapsto \mathbb{N}_0 \cup \{0\}$
    such that if $\exists q, r \in R \quad a = b \cdot q + r$ and $r = 0$ or $a, b \in R \quad F(r) < F(b)$
- Polynomial ring $R[X]$
  - R is a commutative unitary ring
  - set of all polynomials with coefficients $\in R$