Computational Algebra
-
Transcript

Fabio Gratl

April 27, 2015

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 $a_{n-1} \neq$ 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 $c_i := a_i + b_i + \sigma_i \quad$ and $\quad \sigma := 0$

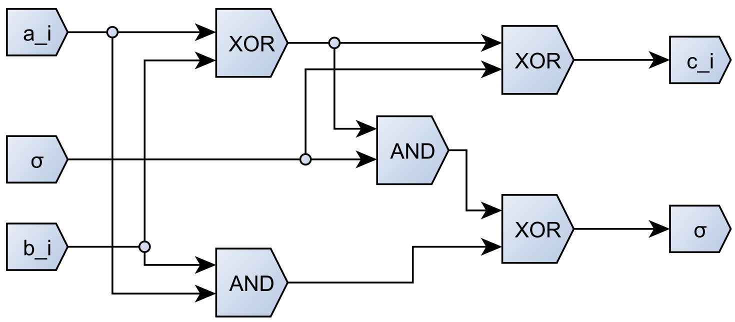
(4) if $(c_i \geq B)$

(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:



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 $n + m$:
$n \cdot m \leq \frac{1}{2}(n+m)^2 \to O((n+m)^2)$
$\Rightarrow$ 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:

$k = 0 : \Theta(k) = 1$

$$\begin{aligned} k-1 \to k = \Theta(k) &= 3\Theta(k-1) + c2^{k-1} \\ &\leq 3(3^{k-1} + 2c(3^{k-r} - 2^{k-1})) + c2^k \\ &= 3^k + 2c(3^k - 2^k) \end{aligned}$$

So $\Theta(k) \leq (2c+1)3^k$

Now $l(x) \leq n$ hence $2^{k-1} < n$ by minimality of k

So $k - 1 < \log_2 n$

$$\begin{aligned} \Rightarrow \Theta(k) &\leq 3(2c+1)3^{\log_2(n)} \\ &= 3(2c+1)2^{\log_2(3)\log_2(n)} \\ &= 3(2c+1)n^{\log_2(3)} \qquad \square \end{aligned}$$

One can modify the terminal condition of Karasuba 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

$\hat{f}(\omega) = \int_{\mathbb{R}} f(t) e^{-i\omega t} dt$ (if it exists)

Think of ω as frequency.

Definition (Convolution)

Let $f, g : \mathbb{R} \mapsto \mathbb{C}$

$(f * g)(x) = \int_{\mathbb{R}} f(t) g(x-t) dt$

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. $e^{2\pi \frac{i}{n}} \in \mathbb{C}$)

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 $\hat{a}_i = \sum_{j=0}^{n-1} \mu^{ij} a_j$
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 $k = 1 \quad (\Rightarrow \mu = 1) \quad$ 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 $\bar{B}^0 + \bar{3}^1 = \bar{u} \neq \bar{0}$ 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 $\hat{a}_j = \sum_{k=0}^{n-1} \mu^{jk} a_k$
$DFT_{\mu^{-1}}(\hat{a}) = (\hat{\hat{a}}_0, ..., \hat{\hat{a}}_1)$
with $\hat{\hat{a}}_i \sum_{i=0}^{n-1} \mu^{-ij} \sum_{k=0}^{n-i} \mu^{jk} a_k = \sum_{k=0}^{n-1} (a_k \cdot \underbrace{\sum_{i=0}^{n-1} \mu^{j(k-i)}}_{=0 \text{ if } n \neq k-i \ \text{(i.e.} k=i)}) = a_i \cdot n$ $\qquad\square$

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 $h > 0 \vee char(R) \neq 2$
$\to \mu$ 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 $i = 2^{k-s} \cdot r \quad$ with r odd $\vee\ s > 0$

$$\sum_{j=0}^{2^k-1} \mu^{ij} = \sum_{l=0}^{2^{k-s}-1} \sum_{j=0}^{2^s-1} \mu^{i(l \cdot 2^s + j)}$$

$$\mu^{i \cdot 2^s} = 1$$

$$i \cdot 2^s = 2^{k-s} \sum_{j=0}^{2^s-1} \mu^{ij} = 2^{k-s} \sum_{j=0}^{2^{s-1}-1} (\mu^{ij} + \mu^{i(2^{s-1}+j)})$$

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 $\sum_{j=0}^{n-1} \mu^{ij} = 0 \qquad \square$

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 : $f, g \in R[x] \quad$ 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 $\hat{f} = DFT_\mu(f), \ \hat{g} = DFT_\mu(g) \quad$ with $f, g \in R^n$

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

(3) compute $(h_0, ..., h_{n-1}) = DFT_{\mu^{-1}}\hat{h} \quad$ (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 $\bar{x} \in R$ with $0 \leq x \leq 2^l$

- Addition: $x \bar{+} \bar{y}$
  (1) compute $x + y \in \mathbb{N}$: $O(l)$
  (2) if $x + y > 2^l + 1$ subtract $2^l + 1$: $O(l)$
- Multiplication by $\bar{2}^i$ $(0 \leq i < l)$
  (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 $\bar{2}^i$ $(l \leq i < 2l - 1)$
  (1) Multiplication by $\bar{2}^{i-l}$: $O(l)$
  (2) take negative $\bar{2}^i \cdot \bar{x} = -\bar{2}^{i-l} \cdot \bar{x}$: $O(l)$

1.1.18 Proposition 18 (Sort of summary)

Let $k, r \in \mathbb{N}$, $r > 0$, $m = 2^{2^k \cdot r} + 1$, $R = \mathbb{Z}/(m)$, $\mu = \bar{2}^r \in R$
$\Rightarrow 2 \in R$ 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 $B = 2^{2^k}$, $m = 2^{2^k \cdot 4} + 1$, $R = \mathbb{Z}/(m)$, $\mu = \bar{2}^4 \in R$
(⇒ so μ is a good primitive 2^{k+1}-th root of 1)

(4) write $x = \sum_{i=0}^{2^k-1} x_i \cdot B^i$, 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
$\to$ 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 $z := \sum_{j=0}^{2^{k+1}-1} z_i \cdot B^i$

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 $k \geq 4$:
$\Theta(k) \leq 2^{k+1} \cdot \Theta(\frac{k+3}{2}) + c \cdot k \cdot 2^{2 \cdot k}$ 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 rouhly is recursion depth
Have $2^{u-1} \leq k-3 \underbrace{\Rightarrow}_{\text{claim}} \Lambda(k) \leq 16^u \cdot D + c \cdot \underbrace{(k-3)}_{<2^u} \cdot \frac{8^u - 1}{7} + 3c \cdot \frac{16^u - 1}{15} \quad \in O(16^u)$
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 ≥ 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
   ($\rightarrow$ 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)

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

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

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

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

(3) if $r \leq 2^i \cdot b$ 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 $r\mathbb{Z}$ 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: TODO TABBING
$a = 287, \quad b = 126$
$287 = 2 \cdot 126 + 35 \qquad (1)$
$126 = 4 \cdot 35 - 14 \qquad (2)$
$35 = (-2) \cdot (-14) + 7 \qquad (3)$
$-14 = (-2) \cdot 7 + 0 \qquad (4)$
So: $7|(-14) \underbrace{\Rightarrow}_{(3)} 7|35$
$\underbrace{\Rightarrow}_{(2)} 7|126$
$\underbrace{\Rightarrow}_{(1)} 7|287$
On the other hand take a common divisor d; $\quad d|287; \quad d|126$
$\underbrace{\Rightarrow}_{(1)} d|d \underbrace{\Rightarrow}_{(2)} d|14 \underbrace{\Rightarrow}_{(3)} d|7$

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 $\leq c \cdot \underbrace{\sum_{i=1}^{k} n_i \cdot (n_{i-1} - n_i + 1)}_{=:\sigma(n_0,\ldots,n_k)}$

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 $\sigma(n_0, ..., n_k) = m \cdot (n - m + 1) + \sum_{i=2}^{m+1} (m - i + 1) \cdot 2 = m \cdot n - m^2 + m + m(m - 1) = m * n.$
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

polynomial rings measure not abs value but max power

2 Notes

- $a|b$
  a is divisible by b
- $ord(a)$
- $char(A)$ 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
- $\lg(x) = \log_2(x)$
- 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