5 Divide & Conquer

10 November 2025

Prof. Dr. Sebastian Wild

CS566 (Wintersemester 2025–26)
Philippps-Universität Marburg

version 2025-11-17 22:18 H

Learning Outcomes

Unit 5: ***Divide & Conquer***

1. Know the steps of the Divide & Conquer paradigm.
2. Be able to solve simple Divide & Conquer recurrences.
3. Be able to design and analyze new algorithms using the Divide & Conquer paradigm.
4. Know the performance characteristics of selection-by-rank algorithms.
5. Know the divide and conquer approaches for integer multiplication, matrix multiplication, finding majority elements, and the closest-pair-of-points problem.

Outline

5 Divide & Conquer

Divide and conquer

Divide and conquer *idiom* (Latin: *divide et impera*)
to make a group of people disagree and fight with one another
so that they will not join together against one (Merriam-Webster Dictionary)

$\rightsquigarrow$ in politics & algorithms, many independent, small problems are better than one big one!

Divide-and-conquer algorithms:

1. Break problem into smaller, independent subproblems. (Divide!)
2. Recursively solve all subproblems. (Conquer!)
3. Assemble solution for original problem from solutions for subproblems.

Examples:

- Mergesort
- Quicksort
- Binary search
- (arguably) Tower of Hanoi

5.1 Divide & Conquer Recurrences

Back-of-the-envelope analysis

- before working out the details of a D&C idea,
  it is often useful to get a quick indication of the resulting performance
  - don't want to waste time on something that's not competitive in the end anyways!
- since D&C is naturally recursive, running time often not obvious
  instead: given by a recursive equation
- unfortunately, rigorous analysis often tricky
  - Remember mergesort?

$$C(n) = \begin{cases} 0 & n \le 1 \\ C(\lfloor n/2 \rfloor) + C(\lceil n/2 \rceil) + 2n & n \ge 2 \end{cases}$$

$$\rightsquigarrow C(n) = 2n\lfloor \lg(n) \rfloor + 2n - 4 \cdot 2^{\lfloor \lg(n) \rfloor} \text{ 😱}$$

$$= \Theta(n \log n) \text{ 😌}$$

- the following method works for many typical cases to give the right **order of growth**

The Master Method

- Assume a stereotypical D&C algorithm
  - a recursive calls on (for some constant $a > 0$)
  - subproblems of size n/b (for some constant $b > 1$)
  - with non-recursive "conquer" effort $f(n)$ (for some function $f : \mathbb{R} \to \mathbb{R}$)
  - base case effort d (some constant $d > 0$)

$\rightsquigarrow$ running time $T(n)$ satisfies

$$\boxed{T(n) \;=\; \begin{cases} a \cdot T\left(\frac{n}{b}\right) \;+\; f(n) & n > 1 \\ d & n \le 1 \end{cases}}$$

Theorem 5.1 (Master Theorem)

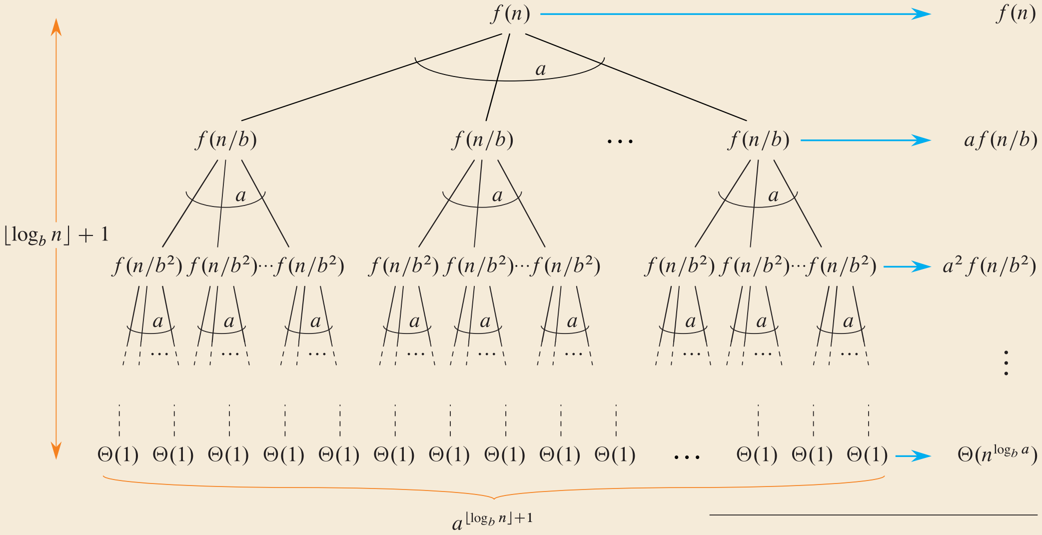
With $c := \log_b(a)$, we have for the above recurrence:

(a) $T(n) = \Theta(n^c)$ if $f(n) = O(n^{c-\varepsilon})$ for constant $\varepsilon > 0$.

(b) $T(n) = \Theta(n^c \log n)$ if $f(n) = \Theta(n^c)$.

(c) $T(n) = \Theta(f(n))$ if $f(n) = \Omega(n^{c+\varepsilon})$ for constant $\varepsilon > 0$ **and** f satisfies the regularity condition $\exists n_0, \alpha < 1\ \forall n \ge n_0 \ :\ a \cdot f\left(\frac{n}{b}\right) \le \alpha f(n)$.

Master Theorem – Intuition & Proof Idea



Total: $\Theta(n^{\log_b a}) + \sum_{j=0}^{\lfloor \log_b n \rfloor} a^j f(n/b^j)$

Figure 4.3 of Cormen et al. *Introduction to Algorithms* 4th ed.

When it's fine to ignore floors and ceilings

Lemma 5.2 (Polynomial-growth master method)

If the toll function $f(n)$ satisfies the *polynomial-growth condition*,
then the Θ-class of the solution of a D&C recurrence remains the same
when ignoring floors and ceilings on subproblem sizes. ◀

The ***polynomial-growth condition***

- $f : \mathbb{R}_{>0} \to \mathbb{R}$ satisfies the *polynomial-growth condition* if

$$\exists n_0 \; \forall C \ge 1 \; \exists D > 1 \quad \forall n \ge n_0 \; \forall c \in [1, C] \;\; : \;\; \frac{1}{D} f(n) \;\le\; f(cn) \;\le\; D f(n)$$

- intuitively: increasing n by up to a factor C (and anywhere in between!)
  changes the function value by at most a factor $D = D(C)$
  (for sufficiently large n)

- examples: $f(n) = \Theta(n^{\alpha} \log^{\beta}(n) \log\log^{\gamma}(n))$ for constants α, β, γ (zero allowed)
  $\rightsquigarrow$ f satisfies the polynomial-growth condition

A Rigorous and Stronger Meta Theorem

Theorem 5.3 (Roura's Discrete Master Theorem)

Let $T(n)$ be recursively defined as

$$T(n) = \begin{cases} b_n & 0 \le n < n_0, \\ f(n) + \sum_{d=1}^{D} a_d \cdot T\left(\frac{n}{b_d} + r_{n,d}\right) & n \ge n_0, \end{cases}$$

where $D \in \mathbb{N}$, $a_d > 0$, $b_d > 1$, for $d = 1, \ldots, D$ are constants, functions $r_{n,d}$ satisfy $|r_{n,d}| = O(1)$ as $n \to \infty$, and function $f(n)$ satisfies $f(n) \sim B \cdot n^\alpha (\ln n)^\gamma$ for constants $B > 0$, α, γ.

Set $H = 1 - \sum_{d=1}^{D} a_d \,(1/b_d)^\alpha$; then we have:

(a) If $H < 0$, then $T(n) = O(n^{\tilde{\alpha}})$, for $\tilde{\alpha}$ the unique value of α that would make $H = 0$.

(b) If $H = 0$ and $\gamma > -1$, then $T(n) \sim f(n)\ln(n)/\tilde{H}$ with constant $\tilde{H} = (\gamma + 1) \sum_{d=1}^{D} a_d\, b_d^{-\alpha} \ln(b_d)$.

(c) If $H = 0$ and $\gamma = -1$, then $T(n) \sim f(n)\ln(n)\ln(\ln(n))/\hat{H}$ with constant $\hat{H} = \sum_{d=1}^{D} a_d\, b_d^{-\alpha} \ln(b_d)$.

(d) If $H = 0$ and $\gamma < -1$, then $T(n) = O(n^\alpha)$.

(e) If $H > 0$, then $T(n) \sim f(n)/H$.

5.2 Order Statistics

Selection by Rank

- Standard data summary of numerical data: (Data scientists, listen up!)
  - mean, standard deviation
  - min/max (range)
  - histograms

  easy to compute in $\Theta(n)$ time
  - median, quartiles, other quantiles (a.k.a. order statistics)

  
  computable in $\Theta(n)$ time?

General form of problem: **Selection by Rank**

- **Given:** array $A[0..n)$ of numbers and number $k \in [0..n)$.
- **Goal:** find element that would be in position k if A was sorted (kth smallest element).

  but 0-based & counting dups

- $k = \lfloor n/2 \rfloor \rightsquigarrow$ median; $k = \lfloor n/4 \rfloor \rightsquigarrow$ lower quartile
  $k = 0 \rightsquigarrow$ minimum; $k = n - \ell \rightsquigarrow$ ℓth largest

Quickselect

- Key observation: Finding the element of rank k seems hard.
  But computing the rank of a given element is easy!
  count smaller elements

$\rightsquigarrow$ Pick any element $A[b]$ and find its rank j.
  - $j = k$? $\rightsquigarrow$ Lucky Duck! Return chosen element and stop
  - $j < k$? $\rightsquigarrow$ ... not done yet. But: The $j + 1$ elements smaller than $\le A[b]$ can be excluded!
  - $j > k$? $\rightsquigarrow$ similarly exclude the $n - j$ elements $\ge A[b]$

- partition function from Quicksort:
  - returns the rank of pivot
  - separates elements into smaller/larger

$\rightsquigarrow$ can use same building blocks

```
procedure quickselect(A[l..r), k):
    if r − l ≤ 1 then return A[l]
    b := choosePivot(A[l..r))
    j := partition(A[l..r), b)
    if j == k
        return A[j]
    else if j < k
        quickselect(A[j + 1..r), k)
    else // j > k
        quickselect(A[l..j), k)
```

Quickselect – Iterative Code

Recursion can be replaced by loop (*tail-recursion elimination*)

```
procedure quickselect(A[l..r), k):
    if r − l ≤ 1 then return A[l]
    b := choosePivot(A[l..r))
    j := partition(A[l..r), b)
    if j == k
        return A[j]
    else if j < k
        quickselect(A[j + 1..r), k)
    else // j > k
        quickselect(A[l..j), k)
```

```
procedure quickselectIterative(A[0..n), k):
    l := 0;  r := n
    while r − l > 1
        b := choosePivot(A[l..r))
        j := partition(A[l..r), b)
        if j ≥ k then r := j − 1
        if j ≤ k then l := j + 1
    return A[k]
```

- implementations should usually prefer iterative version
- analysis more intuitive with recursive version

Quickselect – Analysis

```
procedure quickselect(A[l..r), k):
    if r − l ≤ 1 then return A[l]
    b := choosePivot(A[l..r))
    j := partition(A[l..r), b)
    if j == k
        return A[j]
    else if j < k
        quickselect(A[j + 1..r), k)
    else // j > k
        quickselect(A[l..j), k)
```

- cost = #cmps
- costs depend on n **and** k
- **worst case:** $k = 0$, but always $j = n - 2$
  - $\rightsquigarrow$ each recursive call makes n one smaller at cost $\Theta(n)$
  - $\rightsquigarrow$ $T(n,k) = \Theta(n^2)$ worst case cost

average case:

- let $T(n,k)$ expected cost when we choose a pivot uniformly from $A[0..n)$

$\rightsquigarrow$ formulate recurrence for $T(n,k)$ similar to BST/Quicksort recurrence

$$T(n,k) \;=\; n \;+\; \frac{1}{n}\sum_{r=0}^{n-1}[r = k]\cdot 0 \;+\; [k < r]\cdot T(r,k) \;+\; [k > r]\cdot T(n-r-1, k-r-1)$$

Quickselect – Average Case Analysis

- $T(n,k) \;=\; n \;+\; \dfrac{1}{n}\displaystyle\sum_{r=0}^{n-1} [r=k]\cdot 0 \;+\; [k<r]\cdot T(r,k) \;+\; [k>r]\cdot T(n-r-1, k-r-1)$
- Set $\hat{T}(n) \;=\; \max_{k\in[0..n)} T(n,k)$

$\rightsquigarrow$ $\hat{T}(n) \;\le\; n \;+\; \dfrac{1}{n}\displaystyle\sum_{r=0}^{n-1} \max\bigl\{\hat{T}(r), \hat{T}(n-r-1)\bigr\}$

- analyze hypothetical, worse algorithm:
  if $r \notin [\frac{1}{4}n, \frac{3}{4}n)$, discard partition and repeat with new pivot!

$\rightsquigarrow$ $\hat{T}(n) \;\le\; \tilde{T}(n)$ defined by $\quad \tilde{T}(n) \;\le\; n \;+\; \frac{1}{2}\tilde{T}(n) \;+\; \frac{1}{2}\tilde{T}(\frac{3}{4}n)$

$\rightsquigarrow$ $\tilde{T}(n) \;\le\; 2n \;+\; \tilde{T}(\frac{3}{4}n)$

- Master Theorem Case 3: $\tilde{T}(n) = \Theta(n)$

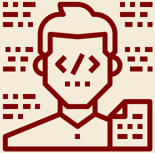
Quickselect Discussion

- $\Theta(n^2)$ worst case (like Quicksort)
- expected cost $\Theta(n)$ (best possible)
- no extra space needed
- adaptations possible to find several order statistics at once
- expected cost can be further improved by choosing pivot from a small sorted sample
  - $\rightsquigarrow$ asymptotically optimal randomized cost: $n + \min\{k, n-k\}$ comparisons in expectation achieved asymptotically by the *Floyd-Rivest algorithm*

5.3 Linear-Time Selection

Interlude – A recurring conversation

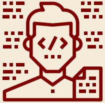
Cast of Characters:



Hi! I'm a *computer science practitioner*.
I love algorithms for the sometimes miraculous **applications** they enable.
I care for things I can **implement** and **that actually work in practice**.

Hi! I'm a *theoretical computer science researcher*.
I find beauty in elegant and **definitive** answers to questions about complexity.
I care for **eternal truths** and mathematically proven facts;
asymptotically optimal is what counts! (Constant factors are secondary.)

Quickselect Disagreements



For practical purposes, (randomized) Quickselect is perfect.

e. g. used in C++ STL `std::nth_element`

Yeah . . . maybe. But can we select by rank in $O(n)$ deterministic **worst case** time?

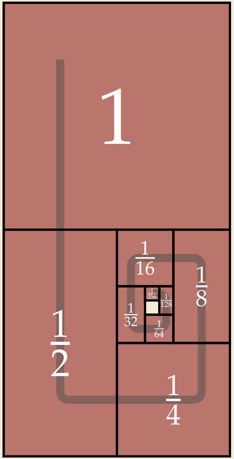
Better Pivots

It turns out, we can!

- All we need is better pivots!
  - If pivot was the exact median,
    we would at least halve #elements in each step
  - Then the total cost of all partitioning steps is $\leq 2n = \Theta(n)$.

But: finding medians is (basically) our original problem!

It totally suffices to find an element of rank αn for $\alpha \in (\varepsilon, 1 - \varepsilon)$
to get overall costs $\Theta(n)$!



The Median-of-Medians Algorithm

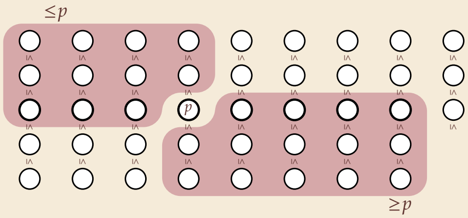
```
procedure choosePivotMoM(A[l..r]):
    m := ⌊n/5⌋
    for i := 0, ..., m − 1
        sort(A[5i..5i + 4])
        // collect median of 5
        Swap A[i] and A[5i + 2]
    return quickselectMoM(A[0..m), ⌊(m−1)/2⌋)

procedure quickselectMoM(A[l..r), k):
    if r − l ≤ 1 then return A[l]
    b := choosePivotMoM(A[l..r))
    j := partition(A[l..r), b)
    if j == k
        return A[j]
    else if j < k
        quickselectMoM(A[j + 1..r), k)
    else // j > k
        quickselectMoM(A[l..j), k)
```

Analysis:

- Note: 2 mutually recursive procedures
  $\rightsquigarrow$ effectively 2 recursive calls!

1. recursive call inside choosePivotMoM on $m \le \frac{n}{5}$ elements
2. recursive call inside quickselectMoM



$\rightsquigarrow$ partition excludes $\sim 3 \cdot \frac{m}{2} \sim \frac{3}{10}n$ elem.

$$\rightsquigarrow \quad C(n) \le \Theta(n) + C(\tfrac{1}{5}n) + C(\tfrac{7}{10}n)$$

$$\le \Theta(n) + C(\tfrac{1}{5}n + \tfrac{7}{10}n)$$

ansatz: overall cost linear

$$= \Theta(n) + C(\tfrac{9}{10}n) \quad \rightsquigarrow \quad C(n) = \Theta(n)$$

5.4 Fast Multiplication

Integer Multiplication

- What's the cost of computing $x \cdot y$ for two integers x and y?

$\rightsquigarrow$ depends on how big the numbers are!

  - If x and y have $O(w)$ bits, multiplication takes $O(1)$ time on word-RAM
  - otherwise, need a dedicated algorithm!

Long multiplication (»Schulmethode«)

- Given $x = \sum_{i=0}^{n-1} x_i 2^i$ and $y = \sum_{i=0}^{n-1} y_i 2^i$, want $z = \sum_{i=0}^{2n-1} z_i 2^i$

```
for i := 0, ..., n − 1
    c := 0
    for j := 0, ..., n − 1
        z_{i+j} := z_{i+j} + c + x_i · y_j
        c := ⌊z_{i+j} / 2⌋
        z_{i+j} := z_{i+j} mod 2
    end for
    z_{i+n} := c
end for
```

- $\Theta(n^2)$ bit operations
- could work with base 2^w instead of 2

  $\rightsquigarrow$ $\Theta\big((n/w)^2\big)$ time
- here: count bit operations for simplicity can be generalized

Example:

easier in binary!

("shift and add")

```
1001010101 * 101101
-------------------
    1001010101
     0000000000
      1001010101
       1001010101
        0000000000
         1001010101
-------------------
    110100011110001
```

Divide & Conquer Multiplication

- assume n is power of 2 (fill up with 0-bits otherwise)
- We can write
  - $x = a_1 2^{n/2} + a_2$ and
  - $y = b_1 2^{n/2} + b_2$
  - for a_1, a_2, b_1, b_2 integers with $n/2$ bits

$$\rightsquigarrow \; x \cdot y \;=\; (a_1 2^{n/2} + a_2) \cdot (b_1 2^{n/2} + b_2) \;=\; \boxed{\boldsymbol{a_1 b_1} 2^n + (\boldsymbol{a_1 b_2} + \boldsymbol{a_2 b_1}) 2^{n/2} + \boldsymbol{a_2 b_2}}$$

  - recursively compute **4** smaller products
  - combine with shifts and additions ($O(n)$ bit operations)

- ...but is this any good?
  - $T(n) \;=\; \mathbf{4} \cdot T(n/2) + \Theta(n)$

  $\rightsquigarrow$ Master Theorem Case 1: $T(n) = \Theta(n^2)$ *...just like the primary school method!?*
    - but Master Theorem gives us a hint: cost is dominated by the leaves

    $\rightsquigarrow$ try to do more work in conquer step!

Karatsuba Multiplication

- how can we do "less divide and more conquer"?

Recall: $x \cdot y \;=\; \boxed{\mathbf{a_1 b_1} 2^n + (\mathbf{a_1 b_2} + \mathbf{a_2 b_1}) 2^{n/2} + \mathbf{a_2 b_2}}$

Let's do some algebra.

$$\begin{aligned} c &:= (a_1 + a_2) \cdot (b_1 + b_2) \\ &= a_1 b_1 + (a_1 b_2 + a_2 b_1) + a_2 b_2 \end{aligned}$$

$\rightsquigarrow$ $(\mathbf{a_1 b_2} + \mathbf{a_2 b_1}) \;=\; c - a_1 b_1 - a_2 b_2$

this can be computed with **3** recursive multiplications

$a_1 + a_2$ and $b_1 + b_2$ still have roughly $n/2$ bits

procedure karatsuba(x, y):
  // Assume x and y are $n = 2^k$ bit integers
  $a_1 := \lfloor x / 2^{n/2} \rfloor$; $a_2 := x \bmod 2^{n/2}$ *// implemented by shifts*
  $b_1 := \lfloor y / 2^{n/2} \rfloor$; $b_2 := y \bmod 2^{n/2}$
  $c_1 :=$ karatsuba(a_1, b_1)
  $c_2 :=$ karatsuba(a_2, b_2)
  $c \;:=$ karatsuba($a_1 + a_2$, $b_1 + b_2$) $- c_1 - c_2$
  return $c_1 2^n + c 2^{n/2} + c_2$ *// shifts and additions*

Analysis:

- nonrecursive cost: only additions and shifts
- all numbers $O(n)$ bits

$\rightsquigarrow$ conquer cost $f(n) = \Theta(n)$

Recurrence:

- $T(n) \;=\; 3T(n/2) + \Theta(n)$
- Master Theorem Case 1

$\rightsquigarrow$ $T(n) = \Theta(n^{\lg 3}) = O(n^{1.585})$

much cheaper (for large n)!

Integer Multiplication

- until 1960, integer multiplication was conjectured to take $\Omega(n^2)$ bit operations

$\rightsquigarrow$ Karatsuba's algorithm was a big breakthrough

  - which he discovered as a student!

- idea can be generalized to breaking numbers into $k \geq 2$ parts (*Toom-Cook algorithm*)

- asymptotically *much* better algorithms are now known!
  - e. g., the *Schönhage-Strassen algorithm* with $O(n \log n \log \log n)$ bit operations (!)
  - these are based on the *Fast Fourier Transform* (FFT) algorithm
    - numbers = polynomials evaluated at base (e. g., $z = 2$)

    $\rightsquigarrow$ multiplication of numbers = convolution of polynomials

    - FFT makes computation of this convolution cheap by computing the polynomial via interpolation
    - Schönhage-Strassen adds careful finite-field algebra to make computations efficient

Matrix Multiplication

- The same trick can also be used for faster matrix multiplication

entry of A in row i and column k

- Recall: For $A, B \in \mathbb{R}^{n\times n}$ we define $C = A \cdot B$ via $c_{i,j} = \sum_{k=1}^{n} a_{i,k}\, b_{k,j}$

$\rightsquigarrow$ Naive cost: n^2 sums with n terms each $\rightsquigarrow$ $\Theta(n^3)$ arithmetic operations

- Can use D&C as follows (assuming n is a power of 2 again)
  - Decompose (cut in half hor. & vert.) $A = \begin{pmatrix} A_{1,1} & A_{1,2} \\ A_{2,1} & A_{2,2} \end{pmatrix}$, $B = \begin{pmatrix} B_{1,1} & B_{1,2} \\ B_{2,1} & B_{2,2} \end{pmatrix}$, $C = \begin{pmatrix} C_{1,1} & C_{1,2} \\ C_{2,1} & C_{2,2} \end{pmatrix}$

  $\rightsquigarrow$ We get C as

$$
\begin{aligned}
C_{1,1} &= A_{1,1} \cdot B_{1,1} + A_{1,2} \cdot B_{2,1} \\
C_{1,2} &= A_{1,1} \cdot B_{1,2} + A_{1,2} \cdot B_{2,2} \\
C_{2,1} &= A_{2,1} \cdot B_{1,1} + A_{2,2} \cdot B_{2,1} \\
C_{2,2} &= A_{2,1} \cdot B_{1,2} + A_{2,2} \cdot B_{2,2}
\end{aligned}
$$

(note "$\cdot$" and "+" operate on matrices here)

4 matrix sums with $(\frac{n}{2})^2$ entries each

  - 8 recursive matrix multiplications on two $\frac{n}{2} \times \frac{n}{2}$ matrices + $\Theta(n^2)$ summations
  - # operations $T(n) = 8T(n/2) + \Theta(n^2)$

  $\rightsquigarrow$ Master Theorem Case 1: $T(n) = \Theta(n^3)$ 😞 (but: still useful for better memory locality!)

Strassen Algorithm for Matrix Multiplication

- Observation (again): Can do more conquer for less divide!
- We recursively compute the following **7** products:

$$\begin{aligned}
M_1 &:= (A_{1,2} - A_{2,2}) \cdot (B_{2,1} + B_{2,2}) \\
M_2 &:= (A_{1,1} + A_{2,2}) \cdot (B_{1,1} + B_{2,2}) \\
M_3 &:= (A_{1,1} - A_{2,1}) \cdot (B_{1,1} + B_{1,2}) \\
M_4 &:= (A_{1,1} + A_{1,2}) \cdot B_{2,2} \\
M_5 &:= A_{1,1} \cdot (B_{1,2} - B_{2,2}) \\
M_6 &:= A_{2,2} \cdot (B_{2,1} - B_{1,1}) \\
M_7 &:= (A_{2,1} + A_{2,2}) \cdot B_{1,1}
\end{aligned}$$

$\rightsquigarrow$ We then obtain the 4 parts of C as

$$\begin{aligned}
C_{1,1} &= M_1 + M_2 - M_4 + M_6 \\
C_{1,2} &= M_4 + M_5 \\
C_{2,1} &= M_6 + M_7 \\
C_{2,2} &= M_2 - M_3 + M_5 - M_7
\end{aligned}$$

(Proof: left as exercise 😉)

Analysis:

- **conquer step:** larger but still $O(1)$ # matrix add/subtract

$\rightsquigarrow$ $\Theta(n^2)$ operations for conquer

$\rightsquigarrow$ total # arithmetic operations $T(n) = \mathbf{7}\,T(n/2) + \Theta(n^2)$

$\rightsquigarrow$ Master Theorem Case 1: $T(n) = \Theta(n^{\lg 7}) = O(n^{2.808})$

Open Problems

*Multiplication is extremely fundamental, but its **computational complexity** is an **open problem** and subject of active research!*

Integer multiplication:

- **conjectured** to require $\Omega(n \log n)$ bit operations (no proof known!)
- Harvey & van der Hoeven **2021**: $O(n \log n)$ algorithm possible!

Matrix multiplication (MM):

- more relevant than it might seem since complexity identical to
  - computing inverse matrices, determinants
  - Gaussian elimination ($\rightsquigarrow$ solving systems of linear equations)
  - recognition of context free languages

$\rightsquigarrow$ best exponent even has standard notation: smallest $\omega \in [2, 3)$ so that MM takes $O(n^{\omega})$ operations

- Big open question: Is $\omega > 2$?
- best known bound: $\omega \leq 2.371339$ (from 2024!)

Timeline of matrix multiplication exponent

Year	Bound on omega	Authors
1969	2.8074	Strassen[1]
1978	2.796	Pan[10]
1979	2.780	Bini, Capovani [it], Romani[11]
1981	2.522	Schönhage[12]
1981	2.517	Romani[13]
1981	2.496	Coppersmith, Winograd[14]
1986	2.479	Strassen[15]
1990	2.3755	Coppersmith, Winograd[16]
2010	2.3737	Stothers[17]
2012	2.3729	Williams[18][19]
2014	2.3728639	Le Gall[20]
2020	2.3728596	Alman, Williams[21][22]
2022	2.371866	Duan, Wu, Zhou[23]
2024	2.371552	Williams, Xu, Xu, and Zhou[2]
2024	2.371339	Alman, Duan, Williams, Xu, Xu, and Zhou[24]

5.5 Majority

Majority

- **Given:** Array $A[0..n)$ of objects
- **Goal:** Check of there is an object x that occurs at $> \frac{n}{2}$ positions in A
  if so, return x
- Naive solution: check each $A[i]$ whether it is a majority $\rightsquigarrow$ $\Theta(n^2)$ time
- Assumption: all we can do to elements is ask "$x = y$?"

Majority – Divide & Conquer

Can be solved faster using a simple Divide & Conquer approach:

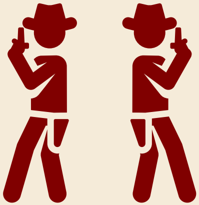
- If A has a majority, that element must also be a majority of at least one half of A.
- $\rightsquigarrow$ Can find majority (if it exists) of left half and right half recursively
- $\rightsquigarrow$ Check these ≤ 2 candidates.
- Costs similar to mergesort: $\Theta(n \log n)$

```
procedure majority(A[0..n)):
    if n == 1 then return A[0] end if
    k := ⌊n/2⌋
    M_ℓ := majority(A[0..k))
    M_r := majority(A[k..n))
    if M_ℓ == M_r then return M_ℓ end if
    m_ℓ := 0;  m_r := 0
    for i := 0, ..., n − 1
        if A[i] == M_ℓ then m_ℓ = m_ℓ + 1 end if
        if A[i] == M_r then m_r = m_r + 1 end if
    end for
    if m_ℓ ≥ k + 1
        return M_ℓ
    else if m_r ≥ k + 1
        return M_r
    else
        return NO_MAJORITY_ELEMENT
```

Majority – Linear Time

We can actually do much better!

```
def MJRTY(A[0..n))
    c := 0
    for i := 1, ..., n − 1
        if c == 0
            x := A[i]; c := 1
        else
            if A[i] == x then c := c + 1 else c := c − 1
    return x
```



- MJRTY($A[0..n)$) returns *candidate* majority element
- either that candidate is the majority element or none exists(!)

Clearly $\Theta(n)$ time

5.6 Closest Pair of Points in the Plane

Closest Pair of Points in the Plane

- **Given:** Array $P[0..n)$ of points in the plane ($\mathbb{R}^2$)
  each has x and y coordinates: $P[i].x$ and $P[i].y$
- **Goal:** Find pair $P[i], P[j]$ that is closest in (Euclidean) distance
  i. e., i and j that minimize $d_2(P[i], P[j]) = \sqrt{(P[i].x - P[j].x)^2 + (P[i].y - P[j].y)^2}$
- Naive solution: compute distance of each pair $\leadsto$ $\Theta(n^2)$ time
  - cost here = # arithmetic operations $\leadsto$ $O(1)$ cost to compute d_2
  - ignore numerical accuracy

    Note: Since $\sqrt{\cdot}$ monotonic, suffices to minimize $d_2{}^2(P[i], P[j])$

$\leadsto$ formally work on the ***real RAM***

  - like word-RAM, but words contain **exact** real numbers
  - support arithmetic operations and comparisons,
    but **not** bitwise operations or $\lfloor \cdot \rfloor$ and $\lceil \cdot \rceil$

- We focus on computing $\delta = \min d_2(P[i], P[j])$
  remembering actual pair of points is an easy modification

Closest Pair – Divide & Conquer

1. Partition points around median x-coordinate $m.x$
2. Recurse on points left resp. right of $m.x$
3. Consider 3 cases of where δ can be achieved:
   a) closest pair left of $m.x$
   b) closest pair right of $m.x$
   c) closest pair straddling $m.x$

Closest Pair – Checking Straddle Pairs

- number of straddle pairs is $\sim \frac{n}{2} \times \frac{n}{2} \quad \rightsquigarrow \quad$ just as slow as brute force!

- **Insight:** Can exclude any points far from dividing line! (cannot be close)
  - precisely: let δ be closest pair distance from (a) and (b)
  - only points with x-coordinate in $m.x \pm \delta$ relevant
  - worst case: no single point excluded!

- **Insight 2:** Also points of vertical distance $> \delta$ cannot be closest!
  - consider points in $m.x \pm \delta$ strip in order sorted by y-coordinate
  - use vertical "sweep lines" and compare only all pairs in $2\delta \times \delta$ rectangle.
  - ... how many points can be in one rectangle?
  - since in left and right subproblem closest dist $\geq \delta$: at most 8.

$\rightsquigarrow$ After sorting by y-coordinate, only do a linear number of distance checks!

Closest Pair – Divide and Conquer is not all

$\rightsquigarrow$ Total running time $T(n) = 2T(\frac{n}{2}) + \Theta(n \log n)$

- Recursion tree method: $T(n) = O(n \log^{2}(n))$
  Roura's Master Theorem shows $T(n) = \Theta(n \log^2 n)$

- Can we do better?

- non-recursive cost is dominated by sorting
  - linear number of straddling pairs of distances to consider
  - median by x-coordinate can be found in linear time (median-of-medians algorithm)!

- **Insight 3:** We sort points **once** at beginning and use stable partitioning.
  - $\rightsquigarrow$ Remain sorted for recursive subproblems $\rightsquigarrow$ no need to sort in conquer step!
  - By also sorting (a copy/pointers) by x-coordinate initially, we can avoid selection algorithm!

Closest Pair – Code

```
procedure closestDist(P[0..N), byX[0..n), byY[0..n)):
    // P contains all N ≥ n points
    // P[byX[0]].x ≤ P[byX[1]].x ≤ ··· ≤ P[byX[n]].x
    // P[byY[0]].y ≤ P[byY[1]].y ≤ ··· ≤ P[byY[n]].y
    if n == 2 return d_2(P[byX[0]], P[byX[1]])
    if n == 3 return min{d_2(P[byX[0]], P[byX[1]]),
                         d_2(P[byX[1]], P[byX[2]]),
                         d_2(P[byX[0]], P[byX[2]])}
    // 1. Split by median x and recurse
    k := ⌊n/2⌋;
    m := P[byX[k]]
    byX_L := byX[0..k);  byX_R := byX[k..n)
    byY_L, byY_R := new empty array list
    for i := 0, ..., n − 1
        if P[byY[i]].x ≤ m.x // breaking ties as in byX
            byY_L.append(byY[i])
        else
            byY_R.append(byY[i])
        end if
    end for
    // ...
```

```
    // ... closestDist continued
    δ_L := closestDist(P, byX_L, byY_L)
    δ_R := closestDist(P, byX_R, byY_R)
    δ := min{δ_L, δ_R}
    // 2. Check straddling pairs
    // Find points close to dividing line
    for i := 0, ..., n − 1
        if |P[byY[i]].x − m.x| ≤ δ
            C.append(byY[i])
        end if
    end for
    // Distance ≤ δ implies within 8 positions in C
    for i := 0, ..., C.size()
        for j := i + 1, ..., i + 7
            δ := min{δ, d_2(P[C[i]], P[C[j]])}
        end for
    end for
    return δ

procedure d_2(P, Q):
    return sqrt((P.x − Q.x)^2 + (P.y − Q.y)^2)
```

Closest Pair – Analysis

- initial sorting of the points: $\Theta(n \log n)$
- time for closestDist fulfills recurrence $T(n) = 2T(\frac{n}{2}) + \Theta(n)$
  - $\rightsquigarrow$ Master Theorem Case 2: $T(n) = \Theta(n \log n)$

$\rightsquigarrow$ Total time $\Theta(n \log n)$