5.4.1 The birthday paradox

- How many people must there be in a room before there is a 50% chance that two of them were born on the same day of the year?

- The answer is surprisingly few
- The paradox is that it is in fact far fewer
  – than the number of days in a year, or
  – even half the number of days in a year

An analysis using indicator random variables

- We use indicator random variables to provide a simple but approximate analysis of the birthday paradox
- For each pair \((i, j)\) of the \(k\) people in the room, define the indicator random variable \(X_{ij}\), for \(1 \leq i < j \leq k\), by
  \[
  X_{ij} = \begin{cases} 
  1 & \text{i and j have the same birthday} \\
  0 & \text{otherwise}
  \end{cases}
  \]
• Once birthday $b_i$ for $i$ is chosen, the probability that $b_j$ is chosen to be the same day is $1/n$, where $n = 365$

$$E[X_{ij}] = \Pr\{i \text{ and } j \text{ have the same birthday} \} = 1/n$$

• Let $X$ be a random variable counting the number of pairs of individuals having the same birthday

$$X = \sum_{i=1}^{k} \sum_{j=i+1}^{k} X_{ij}$$

• Taking expectations of both sides and applying linearity of expectation, we obtain

$$E[X] = E \left[ \sum_{i=1}^{k} \sum_{j=i+1}^{k} X_{ij} \right] = \sum_{i=1}^{k} \sum_{j=i+1}^{k} E[X_{ij}]$$

$$= \binom{k}{2} \frac{1}{n} = \frac{k(k-1)}{2n}$$

• When $k(k-1) \geq 2n$, the expected number of pairs of people with the same birthday is at least 1.
Thus, if we have at least $\sqrt{2n} + 1$ individuals in a room, we can expect at least two to have the same birthday.

For $n = 365$, if $k = 28$, the expected number of pairs with the same birthday is

$$\frac{(28 \cdot 27)}{(2 \cdot 365)} \approx 1.0356$$

With at least 28 people, we expect to find at least one matching pair of birthdays.

Mars has $n = 687$ days, need $k = 38$ aliens.

Analysis using only probabilities gives a different exact number of people, but same asymptotically: $\Theta(\sqrt{n})$.

5.4.2 Balls and bins

Consider tossing identical balls randomly into $b$ bins, numbered $1, 2, \ldots, b$.

Tosses are independent, and on each toss the ball is equally likely to end up in any bin.

The probability that a tossed ball lands in any given bin is $1/b$.

The ball-tossing process is a sequence of Bernoulli trials with a probability $1/b$ of success $\equiv$ the ball falls in the given bin.
• **How many balls fall in a given bin?**
  
  – The number of balls that fall in a given bin follows the binomial distribution $b(k; n, 1/b)$.
  
  – If we toss $n$ balls, the expected number of balls that fall in the given bin is $n/b$.

• **How many balls must we toss, on the average, until a given bin contains a ball?**
  
  – The number of tosses until the given bin receives a ball follows the geometric distribution with probability $1/b$ and
  
  – the expected number of tosses until success is $1/(1/b) = b$.

• **How many balls must we toss until every bin contains at least one ball?**
  
  – Call a toss in which a ball falls into an empty bin a “hit”;
  
  – We want to know the expected number $n$ of tosses required to get $b$ hits;
  
  – We can partition the $n$ tosses into stages;
  
  – The $i$th stage consists of the tosses after the $(i-1)$st hit until the $i$th hit;
  
  – The first stage consists of the first toss, since we are guaranteed to have a hit when all bins are empty.
During the $i^{th}$ stage, $i - 1$ bins contain balls and $b - i + 1$ bins are empty.

For each toss in the $i^{th}$ stage, the probability of obtaining a hit is $(b - i + 1)/b$.

$n_i$ is the number of tosses in the $i^{th}$ stage.

The number of tosses required to get $b$ hits is $n = \sum_{i=1}^{b} n_i$.

Each $n_i$ has a geometric distribution with probability of success $(b - i + 1)/b$.

$$E[n_i] = \frac{b}{b - i + 1}$$

$$E[n] = E\left[\sum_{i=1}^{b} n_i\right] = \sum_{i=1}^{b} E[n_i]$$

$$= \sum_{i=1}^{b} \frac{b}{b - i + 1}$$

$$= b \sum_{i=1}^{b} \frac{1}{i}$$

$$= b \left(\ln b + O(1)\right)$$

By harmonic series.
• It therefore takes approximately $b \ln b$ tosses before we can expect that every bin has a ball.
• This problem is also known as the **coupon collector’s problem**, which says that a person trying to collect each of $b$ different coupons expects to acquire approximately $b \ln b$ randomly obtained coupons in order to succeed.
9 Medians and Order Statistics

• The $i$th order statistic of a set of $n$ elements is the $i$th smallest element
  – E.g., the minimum of a set of elements is the first order statistic ($i = 1$), and the maximum is the $n$th order statistic ($i = n$)
• A median is the “halfway point” of the set
• When $n$ is odd, the median is unique, occurring at $i = (n + 1)/2$
  • When $n$ is even, there are two medians, occurring at $i = n/2$ and $i = n/2 + 1$
  • Thus, regardless of the parity of $n$, medians occur at
    – $[i = (n + 1)/2]$ (the lower median) and
    – $[i = (n + 1)/2]$ (the upper median)
• For simplicity, we use “the median” to refer to the lower median
The problem of selecting the $i$th order statistic from a set of $n$ distinct numbers

We assume that the set contains distinct numbers

– virtually everything extends to the situation in which a set contains repeated values

We formally specify the problem as follows:

– **Input:** A set $A$ of $n$ (distinct) numbers and an integer $i$, with $1 \leq i \leq n$

– **Output:** The element $x \in A$ that is larger than exactly $i - 1$ other elements of $A$

We can solve the problem in $O(n \log n)$ time by heapsort or merge sort and then simply index the $i$th element in the output array

There are faster algorithms

First, we examine the problem of selecting the minimum and maximum of a set of elements

Then we analyze a practical randomized algorithm that achieves an $O(n)$ expected running time, assuming distinct elements
9.1 Minimum and maximum

- How many comparisons are necessary to determine the minimum of a set of \( n \) elements?
- We can easily obtain an upper bound of \( n - 1 \) comparisons
  - examine each element of the set in turn and keep track of the smallest element seen so far.
- In the following procedure, we assume that the set resides in array \( A \), where \( A.length = n \)

\[
\text{MINIMUM}(A) \\
1. \quad \text{min} \leftarrow A[1] \\
2. \quad \text{for } i \leftarrow 2 \text{ to } A.length \\
3. \quad \text{if } \text{min} > A[i] \\
4. \quad \text{min} \leftarrow A[i] \\
5. \quad \text{return } \text{min}
\]

- We can, of course, find the maximum with \( n - 1 \) comparisons as well
This is the best we can do, since we can obtain a lower bound of $n - 1$ comparisons:
- Think of any algorithm that determines the minimum as a tournament among the elements.
- Each comparison is a match in the tournament in which the smaller of the two elements wins.
- Observing that every element except the winner must lose at least one match, we conclude that $n - 1$ comparisons are necessary to determine the minimum.

Hence, the algorithm MINIMUM is optimal w.r.t. the number of comparisons performed.

Simultaneous minimum and maximum

- Sometimes, we must find both the minimum and the maximum of a set of $n$ elements.
- For example, a graphics program may need to scale a set of $(x, y)$ data to fit onto a rectangular display screen or other graphical output device.
- To do so, the program must first determine the minimum and maximum value of each coordinate.
\( \Theta(n) \) comparisons is asymptotically optimal:

- Simply find the minimum and maximum independently, using \( n - 1 \) comparisons for each, for a total of \( 2n - 2 \) comparisons.

- In fact, we can find both the minimum and the maximum using at most \( 3 \lceil n/2 \rceil \) comparisons by maintaining both the minimum and maximum elements seen thus far.

- Rather than processing each element of the input by comparing it against the current minimum and maximum, we process elements in pairs.

- Compare pairs of input elements first with each other, and then we compare the smaller with the current \( \text{min} \) and the larger to the current \( \text{max} \), at a cost of 3 comparisons for every 2 elements.

- If \( n \) is odd, we set both the \( \text{min} \) and \( \text{max} \) to the value of the first element, and then we process the rest of the elements in pairs.

- If \( n \) is even, we perform 1 comparison on the first 2 elements to determine the initial values of the \( \text{min} \) and \( \text{max} \), and then process the rest of the elements in pairs as in the case for odd \( n \).
• If $n$ is odd, then we perform $3\lfloor n/2 \rfloor$ comparisons

• If $n$ is even, we perform 1 initial comparison followed by $3(n-2)/2$ comparisons, for a total of $3n/2 - 2$

• Thus, in either case, the total number of comparisons is at most $3\lfloor n/2 \rfloor$

9.2 Selection in expected linear time

- The selection problem appears more difficult than finding a minimum, but the asymptotic running time for both is the same: $\Theta(n)$
- A divide-and-conquer algorithm $\text{RANDOMIZED-SELECT}$ is modeled after the quicksort algorithm
- Unlike quicksort, $\text{RANDOMIZED-SELECT}$ works on only one side of the partition
- Whereas quicksort has an expected running time of $\Theta(n \log n)$, the expected running time of $\text{RANDOMIZED-SELECT}$ is $\Theta(n)$, assuming that the elements are distinct
• RANDOMIZED-SELECT uses the procedure
    RANDOMIZED-PARTITION of RANDOMIZED-
    QUICKSORT

\textbf{RANDOMIZED-PARTITION}(A, p, r)

1. \( i \leftarrow \text{RANDOM}(p, r) \)
2. exchange \( A[r] \) with \( A[i] \)
3. \textbf{return} \textsc{Partition}(A, p, r)

\textbf{Partitioning the array}

\textbf{PARTITION}(A, p, r)

1. \( x \leftarrow A[r] \)
2. \( i \leftarrow p - 1 \)
3. \textbf{for} \( j \leftarrow p \textbf{ to } r - 1 \)
4. \textbf{if} \( A[j] \leq x \)
5. \( i \leftarrow i + 1 \)
7. exchange \( A[i + 1] \) with \( A[r] \)
8. \textbf{return} \( i + 1 \)
PARTITION always selects an element $x = A[r]$ as a pivot element around which to partition the subarray $A[p..r]$.

As the procedure runs, it partitions the array into four (possibly empty) regions.

At the start of each iteration of the for loop in lines 3–6, the regions satisfy properties, shown above.
At the beginning of each iteration of the loop of lines 3–6, for any array index \( k \),
1. If \( p \leq k \leq i \), then \( A[k] \leq x \)
2. If \( i + 1 \leq k \leq j - 1 \), then \( A[k] > x \)
3. If \( k = r \), then \( A[k] = x \)

Indices between \( j \) and \( r - 1 \) are not covered by any case, and the values in these entries have no particular relationship to the pivot \( x \)

The running time of \textsc{Partition} on the subarray \( A[p..r] \) is \( \Theta(n) \), \( n = p - r + 1 \)

Return the \( i \)th smallest element of \( A[p..r] \)

\textsc{Randomized-Select}(\( A, p, r, i \))
1. if \( p = r \)
2. return \( A[p] \)
3. \( q \leftarrow \text{Randomized-Partition}(A, p, r) \)
4. \( k \leftarrow q - p + 1 \)
5. if \( i = k \) // the pivot value is the answer
6. return \( A[q] \)
7. elseif \( i < k \)
8. return \( \text{Randomized-Select}(A, p, q - 1, i) \)
9. else return \( \text{Randomized-Select}(A, q + 1, r, i - k) \)
• (1) checks for the base case of the recursion
• Otherwise, RANDOMIZED-PARTITION partitions $A[p..r]$ into two (possibly empty) subarrays $A[p..q-1]$ and $A[q+1..r]$ s.t. each element in the former is $\leq A[q] < $ each element of the latter
• (4) computes the # $k$ of elements in $A[p..q]$
• (5) checks if $A[q]$ is $i$th smallest element
• Otherwise, determine in which of the two subarrays the $i$th smallest element lies
• If $i < k$, then the desired element lies on the low side of the partition, and (8) recursively selects it

• If $i > k$, then the desired element lies on the high side of the partition
• Since we already know $k$ values that are smaller than the $i$th smallest element of $A[p..r]$ the desired element is the $(i-k)$th smallest element of $A[q+1..r]$, which (9) finds recursively
• Worst-case running time for RANDOMIZED-SELECT is $\Theta(n^2)$, even to find the minimum
• The algorithm has a linear expected running time, though
• Because it is randomized, no particular input elicits the worst-case behavior
• Let the running time of RANDOMIZED-SELECT on an input array $A[p..r]$ of $n$ elements be a random variable $T(n)$
• We obtain an upper bound on $E[T(n)]$ as follows

• The procedure RANDOMIZED-PARTITION is equally likely to return any element as the pivot
• Therefore, for each $k$ such that $1 \leq k \leq n$, the subarray $A[p..q]$ has $k$ elements (all less than or equal to the pivot) with probability $1/n$
• For $k = 1, 2, ..., n$, define indicator random variables $X_k$ where
  
  $X_k = I\{\text{the subarray } A[p..q] \text{ has exactly } k \text{ elements}\}$
• Assuming that the elements are distinct, we have $E[X_k] = 1/n$

• When we choose $A[q]$ as the pivot element, we do not know, \textit{a priori},
  1) if we will terminate immediately with the correct answer,
  2) recurse on the subarray $A[p \cdots q - 1]$, or
  3) recurse on the subarray $A[q + 1 \cdots r]$

• This decision depends on where the $i$th smallest element falls relative to $A[q]$

• Assuming $T(n)$ to be monotonically increasing, we can upper-bound the time needed for a recursive call by that on largest possible input

• To obtain an upper bound, we assume that the $i$th element is always on the side of the partition with the greater number of elements

• For a given call of \textsc{Randomized-Select}, the indicator random variable $X_k$ has value 1 for exactly one value of $k$, and it is 0 for all other $k$

• When $X_k = 1$, the two subarrays on which we might recurse have sizes $k - 1$ and $n - k$
We have the recurrence

\[ T(n) \leq \sum_{i=1}^{n} X_k \cdot (T(\max(k-1,n-k)) + O(n)) \]

\[ = \sum_{i=1}^{n} X_k \cdot T(\max(k-1,n-k)) + O(n) \]

Taking expected values, we have

\[ E[T(n)] \leq E \left[ \sum_{i=1}^{n} X_k \cdot T(\max(k-1,n-k)) + O(n) \right] \]

\[ = \sum_{i=1}^{n} E[X_k \cdot T(\max(k-1,n-k))] + O(n) \]

The first Eq. on this slide follows by independence of random variables \( X_k \) and \( \max(k-1,n-k) \)

Consider the expression

\[ \max(k-1,n-k) = \begin{cases} k-1 & \text{if } k > \lfloor n/2 \rfloor \\ n-k & \text{if } k \leq \lfloor n/2 \rfloor \end{cases} \]
If \( n \) is even, each term from \( T([n/2]) \) up to \( T(n - 1) \) appears exactly twice in the summation, and if \( n \) is odd, all these terms appear twice and \( T([n/2]) \) appears once.

Thus, we have

\[
E[T(n)] \leq \frac{2}{n} \sum_{k=[n/2]}^{n-1} E[T(k)] + O(n)
\]

We can show that \( E[T(n)] = O(n) \) by substitution.

In summary, we can find any order statistic, and in particular the median, in expected linear time, assuming that the elements are distinct.
Dynamic sets

- Sets are fundamental to computer science
- Algorithms may require several different types of operations to be performed on sets
- For example, many algorithms need only the ability to
  - insert elements into, delete elements from, and test membership in a set
- We call a dynamic set that supports these operations a dictionary

Operations on dynamic sets

- Operations on a dynamic set can be grouped into **queries** and **modifying operations**

**SEARCH**(S, k)
- Given a set S and a key value k, return a pointer x to an element in S such that x.key = k, or NIL if no such element belongs to S

**INSERT**(S, x)
- Augment the set S with the element pointed to by x. We assume that any attributes in element x have already been initialized

**DELETE**(S, x)
- Given a pointer x to an element in the set S, remove x from S. Note that this operation takes a pointer to an element x, not a key value

**MINIMUM**(S)
- A query on a totally ordered set S that returns a pointer to the element of S with the smallest key
11 Hash Tables

- Often one only needs the dictionary operations \textsc{Insert}, \textsc{Search}, and \textsc{Delete}
- Hash table effectively implements dictionaries
- In the worst case, searching for an element in a hash table takes $\Theta(n)$ time
- In practice, hashing performs extremely well
- Under reasonable assumptions, the average time to search for an element is $O(1)$
11.2 Hash tables

- A set $K$ of keys is stored in a dictionary, it is usually much smaller than the universe $U$ of all possible keys
- A hash table requires storage $\Theta(|K|)$ while search for an element in it only takes $O(1)$ time
- The catch is that this bound is for the average-case time

An element with key $k$ is stored in slot $h(k)$; that is, we use a hash function $h$ to compute the slot from the key $k$
- $h: U \rightarrow \{0,1,\ldots,m-1\}$ maps the universe $U$ of keys into the slots of hash table $T[0..m-1]$
- The size $m$ of the hash table is typically $\ll |U|$
- We say that an element with key $k$ hashes to slot $h(k)$ or that $h(k)$ is the hash value of $k$
- If $k_1 \neq k_2$ hash to same slot we have a collision
- Effective techniques resolve the conflict
The ideal solution avoids collisions altogether

We might try to achieve this goal by choosing a suitable hash function $h$

One idea is to make $h$ appear to be random, thus minimizing the number of collisions

Of course, $h$ must be deterministic so that a key $k$ always produces the same output $h(k)$

Because $|U| > m$, there must be at least two keys that have the same hash value;
  - avoiding collisions altogether is therefore impossible
Collision resolution by chaining

- In chaining, we place all the elements that hash to the same slot into the same linked list.
- Slot $j$ contains a pointer to the head of the list of all stored elements that hash to $j$.
- If no such elements exist, slot $j$ contains NIL.
- The dictionary operations on a hash table $T$ are easy to implement when collisions are resolved by chaining.

CHAINED-HASH-INSERT($T, x$)
1. insert $x$ at the head of list $T[h(x.key)]$

CHAINED-HASH-SEARCH($T, k$)
1. search for element with key $k$ in list $T[h(k)]$

CHAINED-HASH-DELETE($T, x$)
1. delete $x$ from the list $T[h(x.key)]$
• Worst-case running time of insertion is $O(1)$
• It is fast in part because it assumes that the element $x$ being inserted is not already present in the table
• For searching, the worst-case running time is proportional to the length of the list; we analyze this operation more closely soon
• We can delete an element (given a pointer) in $O(1)$ time if the lists are doubly linked
• In singly linked lists, to delete $x$, we would first have to find $x$ in the list $T[h(x.key)]$

Analysis of hashing with chaining

• A hash table $T$ of $m$ slots stores $n$ elements, define the load factor $\alpha = n/m$, i.e., the average number of elements in a chain
• Our analysis will be in terms of $\alpha$, which can be $< 1$, $= 1$, or $> 1$

• In the worst-case all $n$ keys hash to one slot
• The worst-case time for searching is thus $\Theta(n)$ plus the time to compute the hash function
• Average-case performance of hashing depends on how well $h$ distributes the set of keys to be stored among the $m$ slots, on the average.

• **Simple uniform hashing** (SUH):
  – Assume that an element is equally likely to hash into any of the $m$ slots, independently of where any other element has hashed to.

• For $j = 0, 1, \ldots, m - 1$, let us denote the length of the list $T[j]$ by $n_j$, so $n = n_0 + \cdots + n_{m-1}$ and the expected value of $n_j$ is

$$E[n_j] = \alpha = n/m$$

**Theorem 11.1** *In a hash table which resolves collisions by chaining, an unsuccessful search takes average-case time $\Theta(1 + \alpha)$, under SUH assumption.*

**Proof** Under SUH, $k$ not already in the table is equally likely to hash to any of the $m$ slots.

The time to search unsuccessfully for $k$ is the expected time to go through list $T[h(k)]$, which has expected length $E[n_{h(k)}] = \alpha$.

Thus, the expected number of elements examined is $\alpha$, and the total time required (incl. computing $h(k)$) is $\Theta(1 + \alpha)$. ■
Theorem 11.2 \hspace{1cm} In a hash table which resolves collisions by chaining, a \textbf{successful} search takes average-case time $\Theta(1 + \alpha)$, under SUH.

Proof \hspace{1cm} We assume that the element being searched for is equally likely to be any of the $n$ elements stored in the table. The number of elements examined during the search for $x$ is one more than the number of elements that appear before $x$ in $x$’s list. Elements before $x$ in the list were all inserted after $x$ was inserted.

- Let us take the average (over the $n$ table elements $x$) of the expected number of elements added to $x$’s list after $x$ was added to the list + 1
- Let $x_i$, $i = 1, 2, \ldots, n$, denote the $i$th element inserted into the table and let $k_i = x_i.key$
- For keys $k_i$ and $k_j$, define the indicator random variable $X_{ij} = 1\{h(k_i) = h(k_j)\}$
- Under SUH, $Pr\{h(k_i) = h(k_j)\} = 1/m$, and by Lemma 5.1, $E[X_{ij}] = 1/m$
Thus, the expected number of elements examined in successful search is

\[
E \left[ \frac{1}{n} \sum_{i=1}^{n} \left( 1 + \sum_{j=i+1}^{n} X_{ij} \right) \right]
\]

\[
= \frac{1}{n} \sum_{i=1}^{n} \left( 1 + \sum_{j=i+1}^{n} E[X_{ij}] \right)
\]

\[
= \frac{1}{n} \sum_{i=1}^{n} \left( 1 + \sum_{j=i+1}^{n} \frac{1}{m} \right)
\]

\[
= 1 + \frac{1}{nm} \sum_{i=1}^{n} (n - i)
\]

\[
= 1 + \frac{1}{nm} \left( \sum_{i=1}^{n} n - \sum_{i=1}^{n} i \right)
\]

\[
= 1 + \frac{1}{nm} \left( n^2 - \frac{n(n+1)}{2} \right)
\]

\[
= 1 + \frac{n-1}{2m} = 1 + \frac{\alpha}{2} + \frac{\alpha}{2n}
\]

Thus, the total time required for a successful search is \(\Theta(2 + \alpha/2 + \alpha/2n) = \Theta(1 + \alpha)\)
If the number of hash-table slots is at least proportional to the number of elements in the table, we have \( n = O(m) \) and, consequently, 
\[
\alpha = \frac{n}{m} = \frac{O(m)}{m} = O(1)
\]

Thus, searching takes constant time on average.

- Insertion takes \( O(1) \) worst-case time.
- Deletion takes \( O(1) \) worst-case time when the lists are doubly linked.
- Hence, we can support all dictionary operations in \( O(1) \) time on average.