Strings and sequences are one of the simplest, most natural and most used forms of storing information.

- natural language, biosequences, programming source code, XML, music, any data stored in a file
- Many algorithmic techniques are first developed for strings and later generalized for more complex types of data such as graphs.

The area of algorithm research focusing on strings is sometimes known as stringology. Characteristic features include

- Huge data sets (document databases, biosequence databases, web crawls, etc.) require efficiency. Linear time and space complexity is the norm.
- Strings come with no explicit structure, but many algorithms discover implicit structures that they can utilize.

Algorithms for spaced alphabet use only character comparisons.

Algorithms for integer alphabet can use more powerful operations such as using a symbol as an address to a table or arithmetic operations to compute a hash function.

Algorithms for constant alphabet can perform almost any operation on characters and even sets of characters in constant time.

The assumption of a constant alphabet in the analysis of an algorithm often indicates one of two things:

- The effect of the alphabet on the complexity is complicated and the constant alphabet assumption is used to simplify the analysis.
- The time or space complexity of the algorithm is heavily (e.g., linearly) dependent on the alphabet size and the algorithm is effectively unusable for large alphabets.

An algorithm is called alphabet-independent if its complexity has no dependence on the alphabet size.

There are many notations for strings.

When describing algorithms, we will typically use the array notation to emphasize that the string is stored in an array:

\[ T = T[0..n] = T[0]T[1]...T[n-1] \]

Note the half-open range notation \([0..n)\) which is often convenient.

In an abstract context, we often use other notations, for example:

- \( a, \beta \in \Sigma^* \)
- \( x = a_1a_2...a_k \) where \( a_i \in \Sigma \) for all \( i \)
- \( w = u v \), \( u, v \in \Sigma^* \) (\( u \) is the concatenation of \( u \) and \( v \))

We will use \(|w|\) to denote the length of a string \( w \).
Individual characters or their positions usually do not matter. The significant entities are the substrings or factors.

**Definition 0.1:** Let \( w = xyz \) for any \( x, y, z \in \Sigma^* \). Then \( x \) is a prefix, \( y \) is a factor (substring), and \( z \) is a suffix of \( w \).

If \( x \) is both a prefix and a suffix of \( w \), then \( x \) is a border of \( w \).

**Example 0.2:** Let \( w = bobobo \). Then
- \( e, b, bobo, boso, bosobobo, bosoobo \) are the prefixes of \( w \)
- \( e, o, bo, obo, sobo, osobo, bosobo \) are the suffixes of \( w \)
- \( e, b, bosobobo \) are the borders of \( w \)
- \( e, o, b, ob, os, ob, bobo, boso, sobo, bosobobo, bosoobo \) are the factors of \( w \).

Note that \( e \) and \( w \) are always suffixes, prefixes, and borders of \( w \). A suffix/prefix/border of \( w \) is proper if it is not \( w \), and nontrivial if it is not \( e \) or \( w \).

A De Bruijn sequence \( B_k \) of order \( k \) for an alphabet \( \Sigma \) of size \( n \) is a cyclic string of length \( n^k \) that contains every string of length \( k \) over the alphabet \( \Sigma \) as a factor exactly once. The cycle can be opened into a string of length \( n^k + k - 1 \) with the same property.

**Example 0.4:** De Bruijn sequences for the alphabet \( \{a, b\} \):
- \( B_2 = aab(a) \)
- \( B_3 = aababbb(\text{aaa}) \)
- \( B_4 = aabaabababbabbbb(\text{aaa}) \)

De Bruijn sequences are not unique. They can be constructed by finding Eulerian or Hamiltonian cycles in a De Bruijn graph.

**Example 0.5:** De Bruijn graph for the alphabet \( \{a, b\} \) that can be used for constructing \( B_2 \) (Hamiltonian cycle) or \( B_3 \) (Eulerian cycle).

An efficient execution of the operations requires that the set is stored as a suitable data structure.

- A (balanced) binary search tree supports the basic operations in \( O(\log n) \) time and range searching in \( O(\log n + r) \) time, where \( n \) is the size of the set and \( r \) is the size of the result.
- An ordered array supports lookup and range searching in the same time as binary search trees. It is simpler, faster and more space efficient in practice, but does not support insertions and deletions.
- A hash table supports the basic operations in constant time (usually randomized) but does not support range queries.

A data structure is called dynamic if it supports insertions and deletions (tree, hash table) and static if not (array). Static data structures are constructed once for the whole set of objects. In the case of an ordered array, this involves another important operation, sorting. Sorting can be done in \( O(n \log n) \) time using comparisons and even faster for integers.

**Trie**

A simple but powerful data structure for a set of strings is the trie. It is a rooted tree with the following properties:

- Edges are labelled with symbols from an alphabet \( \Sigma \).
- For every node \( v \), the edges from \( v \) to its children have different labels.

Each node represents the string obtained by concatenating the symbols on the path from the root to that node.

The trie for a string set \( R \), denoted by \( \text{trie}(R) \), is the smallest trie that has nodes representing all the strings in \( R \). The nodes representing strings in \( R \) may be marked.

**Example 1.1:** \( \text{trie}(R) \) for \( R = \{alice, anna, eliza\} \).

Some Interesting Strings

The Fibonacci strings are defined by the recurrence:

\[
F_0 = \varepsilon \\
F_1 = b \\
F_{i+1} = F_i F_{i-1} \\
\]

where \( F_1 = b \). The infinite Fibonacci string is the limit \( F_\infty \).

For all \( i > 1 \), \( F_i \) is a prefix of \( F_\infty \).

Fibonacci strings have many interesting properties:

- \( |F_i| = F_i \), where \( F_i \) is the \( i \)th Fibonacci number.
- \( F_n \) has exactly \( k+1 \) distinct factors of length \( k \).
- For all \( i > 1 \), we can obtain \( F_i \) from \( F_{i-1} \) by applying the substitutions \( a \rightarrow ab \) and \( b \rightarrow a \) every character.

1. Sets of Strings

Basic operations on a set of objects include:

- **Insert:** Add an object to the set
- **Delete:** Remove an object from the set.
- **Lookup:** Find if a given object is in the set, and if it is, possibly return some data associated with the object.

There can also be more complex queries:

- **Range query:** Find all objects in a given range of values.

There are many other operations too but we will concentrate on these here.

The above time complexities assume that basic operations on the objects including comparisons can be performed in constant time. When the objects are strings, this is no more true:

- The worst case time for a string comparison is the length of the shorter string. Even the average case time for a random set of \( n \) strings is \( O(\log n) \) in many cases, including for basic operations in a balanced binary search tree. We will show an even stronger result for sorting later. And sets of strings are rarely fully random.
- Computing a hash function is slower too. A good hash function depends on all the characters and cannot be computed faster than the length of the string.

For a string set \( R \), there are also new types of queries:

- **Prefix query:** Find all strings in \( R \) that have the query string \( S \) as a prefix. This is a special type of range query.
- **Lcp (longest common prefix) query:** What is the length of the longest prefix of the query string \( S \) that is also a prefix of some string in \( R \).

Thus we need special set data structures and algorithms for strings.

The trie is conceptually simple but it is not simple to implement efficiently. The time and space complexity of a trie depends on the implementation of the child function:

For a node \( v \) and a symbol \( c \in \Sigma \), \( \text{child}(v, c) \) is \( u \) if \( u \) is a child of \( v \) and the edge \( (v, u) \) is labelled with \( c \), and \( \text{child}(v, c) = \perp \) (null) if \( v \) has no such child.

As an example, here is the insertion algorithm:

**Algorithm 1.2:** Insertion into trie

- **Input:** \( \text{trie}(R) \) and a string \( S[0..m] \in R \)

- **Output:** \( \text{trie}(R \cup \{S\}) \)

1. \( v \leftarrow \text{root} \); \( j \leftarrow 0 \)
2. \( \text{while} \ \text{child}(v, S[j]) \neq \perp \ \text{do} \)
3. \( v \leftarrow \text{child}(v, S[j]) \); \( j \leftarrow j + 1 \)
4. \( \text{while} \ j < m \ \text{do} \)
5. \( \text{Create new node } w \) (initializes \( \text{child}(u, c) \) to \( \perp \) for all \( c \in \Sigma \))
6. \( \text{child}(w, S[j]) \leftarrow \perp \)
7. \( v \leftarrow w \); \( j \leftarrow j + 1 \)
8. Mark \( v \) as representative of \( S \)
There are many implementation options for the child function including:

Array: Each node stores an array of size $\sigma$. The space complexity is $O(\sigma N)$, where $N$ is the number of nodes in $\text{trie}(R)$. The time complexity of the child operation is $O(1)$. Requires an integer alphabet.

Binary tree: Replace the array with a binary tree. The space complexity is $O(N)$ and the time complexity $O(\log \sigma)$. Works for an ordered alphabet.

Hash table: One hash table for the whole trie, storing the values $\text{child}(v, c) \neq \bot$. Space complexity $O(N)$, time complexity $O(1)$. Requires an integer alphabet.

A common simplification in the analysis of tries is to assume a constant alphabet. Then the implementation does not matter: Insertion, deletion, lookup and lcp query for a string $S$ take $O(|S|)$ time.

Note that a trie is a complete representation of the strings. There is no need to store the strings separately.

If $R$ is prefix free, the leaves of $\text{trie}(R)$ represent exactly $R$. This simplifies the implementation of the trie.

**Example 1.5:** The trie for $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\}.

![Trie Diagram](image)

The space complexity of a compact trie is $O(|R|)$ (in addition to the strings):

- In a compact trie, every internal node (except possibly the root) has at least two children. In such a tree, there is always at least as many leaves as internal nodes. Thus the number of nodes is at most $2|R|$.
- The edge labels are factors of the input strings. If the input strings are stored separately, the edge labels can be represented in constant space using pointers to the strings.

The time complexities are the same or better than for tries:

- An insertion adds and a deletion removes at most two nodes.
- Lookups may execute fewer calls to the child operation, though the worst case complexity is the same.
- Prefix and range queries are faster even on a constant alphabet (exercise).

**Prefix free string sets**

Many data structures and algorithms for a string set $R$ become simpler if $R$ is prefix free.

**Definition 1.3:** A string set $R$ is prefix free if no string in $R$ is a prefix of another string in $R$.

There is a simple way to make any string set prefix free:

- Let $\$ \notin \Sigma$ be an extra symbol satisfying $\$ < $c$ for all $c \in \Sigma$.
- Append $\$ to the end of every string in $R$.

This has little or no effect on most operations on the set. The length of each string increases by one only, and the additional symbol could be there only virtually.

**Example 1.4:** The set $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\} is not prefix free because $\text{ali}$ is a prefix of $\text{alice}$, but $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\}$ is prefix free.

**Compact Trie**

Tries suffer from a large number of nodes, close to $||R||$ in the worst case.

- For a string set $R$, we use $|R|$ to denote the number of strings in $R$ and $||R||$ to denote the total length of the strings in $R$.

The space requirement can be problematic, since typically each node needs much more space than a single symbol.

Compact tries reduce the number of nodes by replacing branchless path segments with a single edge.

**Example 1.6:** Compact trie for $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\}.

![Compact Trie Diagram](image)

There is also an intermediate form of trie called leaf-path-compact trie, where branchless path segments are compacted only when they end in a leaf.

- Typically (though not in the worst case) this achieves most of the advantages of a compact trie.
- For trie algorithms, this means stopping the normal search, when only one string is remaining in the subtree.

**Example 1.7:** Leaf-path-compact trie for $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\}.

![Leaf-path-compact Trie Diagram](image)

**Ternary Trie**

Tries can be implemented for ordered alphabets but a bit awkwardly using a comparison-based child function. Ternary trie is a simpler data structure based on symbol comparisons.

Ternary trie is like a binary search tree except:

- Each internal node has three children: smaller, equal and larger.
- The branching is based on a single symbol at a given position as in a trie. The position is zero (first symbol) at the root and increases along the middle branches but not along side branches.

There are also compact ternary tries and leaf-path-compated ternary tries based on compacting branchless path segments.

**Example 1.8:** Ternary tries for $\{\text{ali}$, $\text{alice}$, $\text{anna}$, $\text{elias}$, $\text{eliza}$\}.

![Ternary Trie Diagram](image)

Ternary tries have the same asymptotic size as the corresponding ($\sigma$-ary) tries.
A ternary trie is balanced if each left and right subtree contains at most half of the strings in its parent tree.

- The balance can be maintained by rotations similarly to binary search trees.
- We can also get reasonably close to a balance by inserting the strings in the tree in a random order.

Note that there is no restriction on the size of the middle subtree.

Longest Common Prefixes

The standard ordering for strings is the lexicographical order. It is induced by an order over the alphabet. We will use the same symbols \( \leq, <, \geq, \geq \) etc.) for both the alphabet order and the induced lexicographical order.

We can define the lexicographical order using the concept of the longest common prefix.

**Definition 1.9:** The length of the longest common prefix of two strings \( A[0..n] \) and \( B[0..n] \), denoted by \( lcp(A, B) \), is the largest integer \( \ell \leq \min(m, n) \) such that \( A[0..\ell) = B[0..\ell) \).

**Definition 1.10:** Let \( A \) and \( B \) be two strings over an alphabet with a total order \( \preceq \), and let \( \ell = lcp(A, B) \). Then \( A \) is lexicographically smaller than or equal to \( B \), denoted by \( A \preceq B \), if and only if

1. either \( |A| = \ell \)
2. or \( |A| > \ell \) and \( |A[\ell]| < |B[\ell]| \).

A variant of the LCP array sum is sometimes useful.

**Definition 1.13:** For a string \( S \) and a string set \( R \), define
\[
\text{lcp}(S, R) = \max\{|lcp(S, T) | T \in R\}
\]
\[
\Sigma lcp(R) = \sum_{S \subseteq R} lcp(S, R \setminus \{S\})
\]
The relationship of the two measures is shown by the following two results:

**Lemma 1.14:** For \( i \in [2..n] \), \( LCP_{Bi}[i] = lcp(\{S_i, \ldots, S_{i-1}\}) \).

**Lemma 1.15:** \( \Sigma LCP(R) \leq \Sigma lcp(R) \leq 2 \Sigma LCP(R) \).

The proofs are left as an exercise.

The concept of distinguishing prefix is closely related and often used in place of the longest common prefix for sets. The distinguishing prefix of a string is the shortest prefix that separates it from other strings in the set. It is easy to see that \( dp(S, R \setminus S) = lcp(S, R \setminus S) + 1 \) (at least for a prefix free \( R \)).

**Example 1.16:** For \( R = \{ali$, alice$, anna$, elia$, eliza$\} \), \( \Sigma lcp(R) = 13 \) and \( \Sigma dp(R) = 18 \).

String Sorting

\( \Omega(n \log n) \) is a well known lower bound for the number of comparisons needed for sorting a set of \( n \) objects by any comparison based algorithm. This lower bound holds both in the worst case and in the average case.

There are many algorithms that match the lower bound, i.e., sort using \( \Omega(n \log n) \) comparisons (worst or average case). Examples include quicksort, heapsort and mergesort.

If we use one of these algorithms for sorting a set of \( n \) strings, it is clear that the number of symbol comparisons can be more than \( \Omega(n \log n) \) in the worst case. Determining the order of \( A \) and \( B \) needs at least \( lcp(A, B) \) symbol comparisons and \( lcp(A, B) \) can be arbitrarily large in general.

On the other hand, the average number of symbol comparisons for two random strings is \( O(1) \). Does this mean that we can sort a set of random strings in \( O(n \log n) \) time using a standard sorting algorithm?

In a balanced ternary trie each step down either

- moves the position forward (middle branch), or
- halves the number of strings remaining in the subtree (side branch).

Thus, in a balanced ternary trie storing \( n \) strings, any downward traversal following a string \( S \) passes at most \( \log n \) middle edges and at most \( \log n \) side edges.

Thus the time complexity of insertion, deletion, lookup and lcp query is \( O(S + \log n) \).

In comparison based tries, where the child function is implemented using binary search trees, the time complexities could be \( O(S \log \sigma) \), a multiplicative factor \( O(\log n) \) instead of an additive factor \( O(\log n) \).

Prefix and range queries behave similarly (exercise).

An important concept for sets of strings is the LCP (longest common prefix) array and its sum.

**Definition 1.11:** Let \( R = \{S_1, S_2, \ldots, S_n\} \) be a set of strings and assume \( S_1 < S_2 < \cdots < S_n \). Then the LCP array \( LCP_{Bi}[i] \) is defined so that \( LCP_{Bi}[i] = 0 \) and for \( i \in [2..n] \)
\[
LCP_{Bi}[i] = lcp(S_i, S_{i-1})
\]
Furthermore, the LCP array sum is
\[
\Sigma LCP(R) = \sum_{i \in [1..n]} LCP_{Bi}[i]
\]

**Example 1.12:** For \( R = \{ali$, alice$, anna$, elia$, eliza$\} \), \( \Sigma LCP(R) = 7 \) and the LCP array is:

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**Theorem 1.17:** The number of nodes in \( trie(R) \) is exactly \( ||R|| - \Sigma LCP(R) + 1 \), where \( ||R|| \) is the total length of the strings in \( R \).

**Proof.** Consider the construction of \( trie(R) \) by inserting the strings one by one in the lexicographical order using Algorithm 1.2. Initially, the trie has just one node, the root. When inserting a string \( S \), the algorithm executes exactly \( |S| \) rounds of the two while loops, because each round moves one step forward in \( S \). The first loop follows existing edges as long as possible and thus the number of rounds is \( LCP_{Bi}[i] = lcp(S_i, S_{i-1}) \). This leaves \( |S_i| - LCP_{Bi}[i] \) rounds for the second loop, each of which adds one new node to the trie. Thus the total number of nodes in the trie at the end is:
\[
1 + \sum_{i \in [1..n]} |S_i| - LCP_{Bi}[i] = ||R|| - \Sigma LCP(R) + 1
\]

The proof reveals a close connection between \( LCP_{Bi} \) and the structure of the trie. We will later see that \( LCP_{Bi} \) is useful as an actual data structure in its own right.

The following theorem shows that we cannot achieve \( O(n \log n) \) symbol comparisons for any set of \( n \) strings (when \( \sigma = n^{1(\frac{1}{2})} \)).

**Theorem 1.18:** Let \( A \) be an algorithm that sorts a set of objects using only comparisons between the objects. Let \( R = \{S_1, S_2, \ldots, S_n\} \) be a set of \( n \) strings over an ordered alphabet \( \Sigma \) of size \( \sigma \). Sorting \( R \) using \( A \) requires \( \Omega(n \log n \log \sigma) \) symbol comparisons on average, where the average is taken over the initial orders of \( R \).

- If \( \sigma \) is considered to be a constant, the lower bound is \( \Omega(n \log n \log \sigma) \).
- Note that the theorem holds for any comparison based sorting algorithm \( A \) and any string set \( R \). In other words, we can choose \( A \) and \( R \) to minimize the number of comparisons and still not get below the bound.

- Only the initial order is random rather than “any”. Otherwise, we could pick the correct order and use an algorithm that first checks if the order is correct, needing only \( O(n + \Sigma LCP(R)) \) symbol comparisons.

An intuitive explanation for this result is that the comparisons made by a sorting algorithm are not random. In the later stages, the algorithm tends to compare strings that are close to each other in lexicographical order and thus are likely to have long common prefixes.
Proof of Theorem 1.18. Let $k = \lfloor \log_2 n \rfloor / 2$. For any string $\alpha \in \Sigma^k$, let $R_\alpha$ be the set of strings in $R$ having $\alpha$ as a prefix. Let $n_\alpha = |R_\alpha|$. Let us analyze the number of symbol comparisons when comparing strings in $R_\alpha$ against each other.

- Each string comparison needs at least $k$ symbol comparisons.
- No comparison between a string in $R_\alpha$ and a string outside $R_\alpha$ gives any information about the relative order of the strings in $R_\alpha$.
- Thus $A$ needs to do $\Omega(n_\alpha \log n_\alpha)$ string comparisons and $\Omega(k n_\alpha \log n_\alpha)$ symbol comparisons to determine the relative order of the strings in $R_\alpha$.

Thus the total number of symbol comparisons is $\Omega \left( \sum_{\alpha \in \Sigma^k} k n_\alpha \log n_\alpha \right)$ and

$$\sum_{\alpha \in \Sigma^k} k n_\alpha \log n_\alpha \geq k(n - \sqrt{n}) \log \frac{n - \sqrt{n}}{\sqrt{n}} \geq k(n - \sqrt{n}) \log \frac{n - \sqrt{n}}{\sqrt{n} - 1}$$

$$= \Omega(k n \log n) = \Omega(k n \log n_\alpha).$$

Here we have used the facts that $\sigma^k \leq \sqrt{n}$, that $\sum_{\alpha \in \Sigma^k} n_\alpha > n - \sigma^k \geq n - \sqrt{n}$, and that $\sum_{\alpha \in \Sigma^k} n_\alpha \log n_\alpha > (n - \sqrt{n}) \log ((n - \sqrt{n})/\sigma^k)$ (see exercises). □

String Quicksort (Multikey Quicksort)

Quick sort is one of the fastest general purpose sorting algorithms in practice. Here is a variant of quick sort that partitions the input into three parts instead of the usual two parts.

Algorithm 1.20: TernaryQuicksort($R$)

Input: (Multi)set $R$ in arbitrary order.
Output: $R$ in ascending order.

1. if $|R| \leq 1$ then return $R$
2. select a pivot $x \in R$
3. $R_\leq \leftarrow \{ s \in R \mid s < x \}$
4. $R_\geq \leftarrow \{ s \in R \mid s = x \}$
5. $R_\geq \leftarrow \{ s \in R \mid s > x \}$
6. $R_\text{ternary} \leftarrow \text{TernaryQuicksort}(R_\leq)$
7. $R_\text{ternary} \leftarrow \text{TernaryQuicksort}(R_\geq)$
8. return $R_\text{ternary} \cdot R_\leq \cdot R_\geq$

In the initial call, $\ell = 0$.

Example 1.22: A possible partitioning, when $\ell = 2$.

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Theorem 1.23: String quicksort sorts a set $R$ of $n$ strings in $O(\Sigma LCP(R) + n \log n)$ time.

- Thus string quicksort is an optimal symbol comparison based algorithm.
- String quicksort is also fast in practice.

Radix Sort

The $O(n \log n)$ sorting lower bound does not apply to algorithms that use stronger operations than comparisons. A basic example is counting sort for sorting integers.

Algorithm 1.24: CountingSort($R$)

Input: (Multi)set $R = \{ k_1, k_2, \ldots, k_n \}$ of integers from the range $[0, \sigma]$.
Output: $R$ in nondecreasing order in array $J[0..n]$.

1. for $i \leftarrow 0$ to $\sigma - 1$ do $C[i] \leftarrow 0$
2. for $i \leftarrow 1$ to $n$ do $C[k[i]] \leftarrow C[k[i]] + 1$
3. $\sigma \leftarrow 0$
4. for $i \leftarrow 0$ to $\sigma - 1$ do // cumulative sums
5. $\text{tmp} \leftarrow C[i]$; $C[i] \leftarrow \text{sum}$; $\text{sum} \leftarrow \text{sum} + \text{tmp}$
6. for $i \leftarrow 1$ to $n$ do // distribute
7. $J[k[i]] \leftarrow J[k[i]] + C[k[i]]$
8. return $J$

- The time complexity is $O(n + \sigma)$.
- Counting sort is a stable sorting algorithm, i.e., the relative order of equal elements stays the same.
Similarly, the $\Omega(\Sigma LCP(R) + n \log n)$ lower bound does not apply to string sorting algorithms that use stronger operations than symbol comparisons. **Radix sort** is such an algorithm for integer alphabets.

Radix sort was developed for sorting large integers, but it treats an integer as a string of digits, so it is really a string sorting algorithm.

There are two types of radix sorting:

- **MSD radix sort** starts sorting from the beginning of strings (most significant digit).
- **LSD radix sort** starts sorting from the end of strings (least significant digit).

The algorithm assumes that all strings have the same length $m$, but it can be modified to handle strings of different lengths (exercise).

**Theorem 1.27**: LSD radix sort sorts a set $R$ of strings over the alphabet $[0..\sigma]$ in $O(||R|| + n \log n)$ time, where $||R||$ is the total length of the strings in $R$ and $n$ is the longest length of the string in $R$.

**Proof.** Assume all strings have length $m$. The LSD radix sort performs $m$ rounds with each round taking $O(n + \sigma)$ time. The total time is $O(m(n + \sigma)) = O(||R|| + n \log n)$.

The case of variable lengths is left as an exercise. □

- The weakness of LSD radix sort is that it uses $O(||R||)$ time even when $\Sigma LCP(R)$ is much smaller than $||R||$.
- It is best suited for sorting short strings and integers.

**Algorithm 1.29**: MSDRadixSort($R,l$)

**Input**: (Multi)set $R = \{S_1, S_2, \ldots, S_n\}$ of strings over the alphabet $[0..\sigma]$ and the length $l$ of their common prefix.

**Output**: $R$ in ascending lexicographical order.

1. $R_i \leftarrow \{S \in R | |S| = l\}$; $R = R_i \cup R_{l+1}$
2. $R_i \leftarrow \text{CountingSort}(R_i, l)$
3. $(R_0, R_1, \ldots, R_{\sigma-1}) \leftarrow \text{CountingSort}(R_i, l)$
4. for $i \leftarrow 0$ to $\sigma - 1$ do $R_i \leftarrow \text{MSDRadixSort}(R_i, l+1)$
5. return $R_0, R_1, \ldots, R_{\sigma-1}$

Here CountingSort($R_i, l$) not only sorts but also returns the partitioning based on symbols at position $l$. The time complexity is still $O(||R|| + \sigma)$.

The recursive calls eventually lead to a large number of very small sets, but counting sort needs $O(\sigma)$ time no matter how small the set is. To avoid the potentially high cost, the algorithm switches to string quicksort for small sets.

**Lcp-Comparisons**

General (non-string) comparison-based sorting algorithms are not optimal for sorting strings because of an imbalance between effort and result in a string comparison: it can take a lot of time but the result is only a bit of or a trit of useful information.

String quicksort solves this problem by processing the obtained information immediately after each symbol comparison.

An opposite approach is to use a standard comparison with an lcp-comparison, which is the operation LcpCompare($A, B, k$):

- The return value is the pair $(x, \ell)$, where $x \in \{<, =, >\}$ indicates the order, and $\ell = \text{lcp}(A, B)$, the length of the longest common prefix of strings $A$ and $B$.
- The input value $k$ is the length of a known common prefix, i.e., a lower bound on lcp($A, B$). The comparison can skip the first $k$ characters.

Extra time spent in the comparison is balanced by the extra information obtained in the form of the lcp value.

The LSD radix sort algorithm is very simple.

**Algorithm 1.25**: LSDRadixSort($R$)

**Input**: (Multi)set $R = \{S_1, S_2, \ldots, S_n\}$ of strings of length $m$ over alphabet $[0..\sigma]$.

**Output**: $R$ in ascending lexicographical order.

1. for $\ell \leftarrow m - 1$ to 0 do CountingSort($R, \ell$)
2. return $R$

- CountingSort($R, \ell$) sorts the strings in $R$ by the symbols at position $\ell$ using counting sort (with $k_i$ replaced by $S_i[\ell]$). The time complexity is $O(||R|| + \sigma)$.
- The stability of sorting count is essential.

**Example 1.26**: $R = \{\text{cat}, \text{hin}, \text{ham}, \text{bat}\}$.

<p>| | | | | | |</p>
<table>
<thead>
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<tbody>
<tr>
<td>cat</td>
<td>h</td>
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<tr>
<td>ham</td>
<td>b</td>
<td>a</td>
<td>t</td>
<td>t</td>
<td>t</td>
</tr>
<tr>
<td>bat</td>
<td>h</td>
<td>a</td>
<td>t</td>
<td>h</td>
<td>b</td>
</tr>
</tbody>
</table>

It is easy to show that after $i$ rounds, the strings are sorted by suffix of length $i$. Thus, they are fully sorted at the end.

MSD radix sort resembles string quicksort but partitions the strings into $\sigma$ parts instead of three parts.

**Example 1.28**: MSD radix sort partitioning.

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<table>
<thead>
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<tbody>
<tr>
<td>a1</td>
<td>b</td>
<td>a</td>
<td>t</td>
<td>b</td>
<td>l</td>
</tr>
<tr>
<td>a2</td>
<td>g</td>
<td>m</td>
<td>e</td>
<td>n</td>
<td>t</td>
</tr>
<tr>
<td>a3</td>
<td>c</td>
<td>o</td>
<td>a</td>
<td>t</td>
<td>e</td>
</tr>
<tr>
<td>a4</td>
<td>t</td>
<td>e</td>
<td>r</td>
<td>n</td>
<td>a</td>
</tr>
<tr>
<td>a5</td>
<td>l</td>
<td>e</td>
<td>r</td>
<td>n</td>
<td>a</td>
</tr>
</tbody>
</table>

**Theorem 1.30**: MSD radix sort sorts a set $R$ of $n$ strings over the alphabet $[0..\sigma]$ in $O(\Sigma LCP(R) + n \log n)$ time.

**Proof.** Consider a call processing a subset of size $k \geq \sigma$:

- The time excluding the recursive calls but including the call to string quicksort is $O(k)$ and $k$ symbols accessed here will not be accessed again.
- At most $\sigma$ symbols accessed in $S$ will be accessed by the algorithm. Thus the total time spent in this kind of calls is $O(\Sigma LCP(R)) = O(\Sigma LCP(R) + n) = O(\Sigma LCP(R) + n)$.

The calls for a subset of size $k < \sigma$ are handled by string quicksort. Each string is involved in at most one such call. Therefore, the total time over all calls to string quicksort is $O(\Sigma LCP(R) + n \log n)$.

There exists a more efficient variant of MSD radix sort, which uses $\Sigma LCP(R)$ as a lower bound for any algorithm that must access symbols one at a time.

In practice, MSD radix sort is very fast, but it is sensitive to implementation details.

The following result shows how we can use the information from earlier comparisons to obtain a lower bound or even the exact value for an lcp.

**Lemma 1.31**: Let $A$, $B$, and $C$ be strings.

(a) $\text{lcp}(A, C) \geq \min\{\text{lcp}(A, B), \text{lcp}(B, C)\}$.
(b) If $A \leq B \leq C$, then $\text{lcp}(A, C) = \min\{\text{lcp}(A, B), \text{lcp}(B, C)\}$.
(c) If $\text{lcp}(A, B) \neq \text{lcp}(B, C)$, then $\text{lcp}(A, C) = \min\{\text{lcp}(A, B), \text{lcp}(B, C)\}$.

**Proof.** Assume $\ell = \text{lcp}(A, B) \leq \text{lcp}(B, C)$. The opposite case $\text{lcp}(A, B) > \text{lcp}(B, C)$ is symmetric.

(a) Now $A[0..\ell] = B[0..\ell] = C[0..\ell]$ and thus $\text{lcp}(A, C) \geq \ell$.
(b) Either $|A| = \ell$ or $|A| \leq |B| \leq |C|$. In either case, $\text{lcp}(A, C) = \ell$.
(c) Now $\text{lcp}(A, B) < \text{lcp}(B, C)$. If $\text{lcp}(A, C) > \min\{\text{lcp}(A, B), \text{lcp}(B, C)\}$, then $\text{lcp}(A, C) = \min\{\text{lcp}(A, B), \text{lcp}(B, C)\}$, which violates (a).

The above means that the three lcp values between three strings can never be three different values. At least two of them are the same and the third one is the same or bigger.
It can also be possible to determine the order of two strings without comparing them directly.

Lemma 1.32: Let A, B, B′, C be strings such that A ≤ B ≤ C and A ≤ B′ ≤ C.

(a) If lcp(A, B) > lcp(A, B′), then B < B′.
(b) If lcp(B, C) > lcp(B′, C), then B > B′.

Proof. We show (a); (b) is symmetric. Assume to the contrary that B ≥ B′. Then by Lemma 1.31, lcp(A, B) = min(lcp(A, B′), lcp(B′, B)) ≤ lcp(A, B′), which is a contradiction.

Intuitively, the above result makes sense if you think of lcp(·, ·) as a measure of similarity between two strings. The higher the lcp, the closer the two strings are lexicographically.

Algorithm 1.34: StringMerge(\(\mathcal{P}, \mathcal{Q}\))

Input: Sequences \(\mathcal{P} = ((S_1, k_1), \ldots, (S_n, k_n))\) and \(\mathcal{Q} = ((T_1, t_1), \ldots, (T_m, t_m))\)

Output: Merged sequence \(\mathcal{R}\)

1. \(\mathcal{R} \leftarrow \emptyset\)
2. while \(i \leq m\) and \(j \leq n\) do
   3. if \(k_i > t_j\) then \(\mathcal{R} \leftarrow \mathcal{R} \cup \{S_i, k_i\}\) to \(\mathcal{R}\); \(i \leftarrow i + 1\)
   4. else if \(t_j > k_i\) then \(\mathcal{R} \leftarrow \mathcal{R} \cup \{T_j, t_j\}\) to \(\mathcal{R}\); \(j \leftarrow j + 1\)
   5. else \(h \leftarrow \text{lcp}(S_i, T_j)\)
   6. if \(x = \text{lcp}(S_i, T_j)\) then \(\mathcal{R} \leftarrow \mathcal{R} \cup \{S_i, k_i\}\) to \(\mathcal{R}\); \(i \leftarrow i + 1\)
   7. else \(i \leftarrow i + 1\)
   8. append \(\{S_i, k_i\}\) to \(\mathcal{R}\); \(i \leftarrow i + 1\)
   9. \(i \leftarrow i + 1\)
   10. \(j \leftarrow j + 1\)
11. \(j \leftarrow j + 1\)
12. \(k_i \leftarrow h\)
13. \(k_i \leftarrow h\)
14. while \(i \leq m\) do \(\mathcal{R} \leftarrow \mathcal{R} \cup \{S_i, k_i\}\) to \(\mathcal{R}\); \(i \leftarrow i + 1\)
15. while \(j \leq n\) do \(\mathcal{R} \leftarrow \mathcal{R} \cup \{T_j, t_j\}\) to \(\mathcal{R}\); \(j \leftarrow j + 1\)
16. return \(\mathcal{R}\)

Theorem 1.36: String mergesort sorts a set \(\mathcal{R}\) of \(n\) strings in \(O((\Sigma \text{lcp}(\mathcal{R})) + n \log n)\) time.

Proof. If the calls to \(\text{lcp}\)-compare took constant time, the time complexity would be \(O(n \log n)\) by the same argument as with the standard mergesort.

Whenever \(\text{lcp}\)-compare makes more than one, say \(t + 1\) symbol comparisons, one of the lcp values stored with the strings increases by \(t\).

Since the sum of the final lcp values is exactly \(\Sigma \text{lcp}(\mathcal{R})\), the extra time spent in \(\text{lcp}\)-compare is bounded by \(O((\Sigma \text{lcp}(\mathcal{R}))\).

String Mergesort

String mergesort is a string sorting algorithm that uses \(\text{lcp}\)-comparisons. It has the same structure as the standard mergesort: sort the first half and the second half separately, and then merge the results.

Algorithm 1.33: StringMergesort(\(\mathcal{R}\))

Input: Set \(\mathcal{R} = \{S_1, S_2, \ldots, S_n\}\) of strings.

Output: \(\mathcal{R}\) sorted and augmented with \(\text{lcp}\)-compare values.

1. if \(|\mathcal{R}| = 1\) then return \((\mathcal{R}_1, 0)\)

2. \(m \leftarrow \lceil n/2 \rceil\)

3. \(\mathcal{P} \leftarrow \text{StringMergesort}(\{S_1, S_2, \ldots, S_m\})\)
4. \(\mathcal{Q} \leftarrow \text{StringMergesort}(\{S_{m+1}, S_{m+2}, \ldots, S_n\})\)
5. return StringMergesort(\(\mathcal{P}, \mathcal{Q}\))

The output is of the form

\(((T_1, t_1), (T_2, t_2), \ldots, (T_n, t_n))\)

where \(t_i = \text{lcp}(T_i, T_{i-1})\) for \(i > 1\) and \(t_1 = 0\). In other words, \(t_i = \text{lcp}(P_i, P_{i-1})\).

Thus we get not only the order of the strings but also a lot of information about their common prefixes. The procedure StringMergesort uses this information effectively.

Lemma 1.35: StringMerge performs the merging correctly.

Proof. We will show that the following invariant holds at the beginning of each round in the loop on lines (2)–(12):

Let \(X\) be the last string appended to \(\mathcal{R}\) (or \(\varepsilon\) if \(\mathcal{R} = \emptyset\)). Then \(k_i = \text{lcp}(X, S_i)\) and \(t_j = \text{lcp}(X, T_j)\).

The invariant is clearly true in the beginning. We will show that the invariant is maintained and the smaller string is chosen in each round of the loop.

- If \(k_i > t_j\), then \(\text{lcp}(X, S_i) > \text{lcp}(X, T_j)\) and thus \(S_i < T_j\) by Lemma 1.32.
- \(\text{lcp}(S_i, T_j) = \text{lcp}(X, T_j)\) because, by Lemma 1.31, \(\text{lcp}(X, T_j) = \min(\text{lcp}(X, S_i), \text{lcp}(S_i, T_j))\).

Hence, the algorithm chooses the smaller string and maintains the invariant. The case \(t_j > k_i\) is symmetric.

- If \(k_i = t_j\), then clearly \(\text{lcp}(S_i, T_j) \geq k_i\) and the call to \(\text{lcp}\)-compare is safe, and the smaller string is chosen. The update \(t_j \leftarrow h\) or \(k_i \leftarrow h\) maintains the invariant.

String Binary Search

An ordered array is a simple static data structure supporting queries in \(O(\log n)\) time using binary search.

Algorithm 1.37: Binary search

Input: Ordered set \(\mathcal{R} = \{k_1, k_2, \ldots, k_n\}\), query value \(x\).

Output: The number of elements in \(\mathcal{R}\) that are smaller than \(x\).

1. \(\text{left} < 0\) if \(\text{right} = 1\) // output value is in the range [left..right]
2. while \(\text{right} < \text{left} + 1\) do
3. \(\text{mid} \leftarrow \lfloor(\text{left} + \text{right})/2\rfloor\)
4. if \(k_{\text{mid}} < x\) then \(\text{left} \leftarrow \text{mid}\)
5. else \(\text{right} \leftarrow \text{mid}\)
6. return \(\text{left}\)

With strings as elements, however, the query time is

\(O(m \log n)\) in the worst case for a query string of length \(m\)

\(O(\log n \log m)\) on average for a random set of strings.

During the binary search of \(P\) in \(\{S_1, S_2, \ldots, S_n\}\), the basic situation is the following:

- We want to compare \(P\) and \(S_{\text{mid}}\).
- We have already compared \(P\) against \(S_{\text{left}}\) and \(S_{\text{right}}\), and we know that \(S_{\text{left}} \leq P, S_{\text{right}} \geq P\).
- By using \(\text{lcp}\)-comparisons, we know \(\text{lcp}(S_{\text{left}}, P)\) and \(\text{lcp}(P, S_{\text{right}})\).

By Lemmas 1.31 and 1.38,

\(\text{lcp}(P, S_{\text{mid}}) \geq \min(\text{lcp}(S_{\text{left}}, P), \text{lcp}(P, S_{\text{right}}))\)

Thus we can skip \(\min(\text{lcp}(S_{\text{left}}, P), \text{lcp}(P, S_{\text{right}}))\) first characters when comparing \(P\) and \(S_{\text{mid}}\).
**Algorithm 1.39:** String binary search (without precomputed lcps)

**Input:** Ordered string set $R = \{S_1, S_2, \ldots, S_n\}$, query string $P$.

**Output:** The number of strings in $R$ that are smaller than $P$:

1. $left \leftarrow 0$; $right \leftarrow n + 1$
2. $lcp \leftarrow 0$ // $lcp = lcp(S_{left}, P)$
3. $rlcp \leftarrow 0$ // $rlcp = lcp(P, S_{right})$
4. While $right - left > 1$ do
   5. $mid \leftarrow ((left + right)/2)$
   6. $mclp \leftarrow min(lcp, rlcp)$
   7. $(x, mlncp) \leftarrow LcpCompare(P, S_{mid}; mclp)$
   8. If $x = "<"$ then $right \leftarrow mid$; $rlcp \leftarrow mclp$
   9. Else $left \leftarrow mid$; $lcp \leftarrow mlp$
10. Return $left$

- The average case query time is now $O(\log n)$.
- The worst case query time is still $O(m \log n)$ (exercise).

**Lemma 1.40:** Let $A$, $B$, $C$, and $P$ be strings such that $A \leq B \leq C$ and $A \preceq B \preceq C$.

(a) If $lcp(A, B) > lcp(A, B')$, then $B \preceq B'$ and $lcp(B, B') = lcp(A, B')$.
(b) If $lcp(A, B) < lcp(A, B')$, then $B > B'$ and $lcp(B, B') = lcp(A, B)$.
(c) If $lcp(B, C) > lcp(B, B')$, then $B \preceq B'$ and $lcp(B, B') = lcp(B, C)$.
(d) If $lcp(B, C) < lcp(B, B')$, then $B > B'$ and $lcp(B, B') = lcp(B, C)$.
(e) If $lcp(A, B) = lcp(A, C)$ and $lcp(B, C) = lcp(B, C')$, then $lcp(A, B') \geq max(lcp(A, B), lcp(A, C))$.

**Proof:** Cases (a)-(d) are symmetrical, we show (a) $B \preceq B'$ follows from Lemma 1.32. Then by Lemma 1.31, $lcp(A, B') = min(lcp(A, B), lcp(B, B'))$. Since $lcp(A, B) < lcp(A, B')$, we must have $lcp(A, B') = lcp(B, B')$.

In case (e), we use Lemma 1.31:

$lcp(B, B') \geq min(lcp(A, B), lcp(A, B')) = lcp(A, B)$

$lcp(B, B') \geq min(lcp(B, C), lcp(B, C')) = lcp(B, C)$

Thus $lcp(B, B') \geq max(lcp(A, B), lcp(A, C))$.

**Theorem 1.42:** An ordered string set $R = \{S_1, S_2, \ldots, S_n\}$ can be preprocessed in $O(1\text{LCR}(R) + n)$ time and $O(n)$ space so that a binary search with a query string $P$ can be executed in $O(P + \log n)$ time.

**Proof:** The values $LLCP[\text{mid}]$ and $RLCP[\text{mid}]$ can be computed in $O(1\text{LCR}(R) + n)$ time. Thus the arrays $LLCP$ and $RLCP$ can be computed in $O(1\text{LCR}(R) + n)$ time and stored in $O(n)$ space.

The main while loop in Algorithm 1.41 is executed $O(\log n)$ times and everything except LcpCompare on line (11) needs constant time.

If a given LcpCompare call performs $t + 1$ symbol comparisons, $mclp$ increases by at least 1 on line (11). Then on lines (12)–(13), either $lcp$ or $rlcp$ increases by at least $t$, since $mclp = max(lcp, rlcp)$ before LcpCompare. Since $lcp$ and $rlcp$ never decrease and never grow larger than $|P|$, the total number of extra symbol comparisons in LcpCompare during the binary search is $O(|P|)$.

Other comparison-based data structures such as binary search trees can be augmented with lcp information in the same way (study groups).

Any good hash function must depend on all characters. Thus computing $H(S)$ needs $O(|S|)$ time, which can defeat the advantages of hashing:

- A plain comparison of two strings is faster than computing the hashes.
- The main strength of hash tables is the support for constant time insertions and deletions, but inserting a string $S$ into a hash table needs $O(|S|)$ time when the hash computation time is included. Compare this to the $O(1)$ time for a trie under a constant alphabet and the $O(|S| + \log n)$ time for a ternary trie.

However, a hash table can still be competitive in practice. Furthermore, there are situations, where a full computation of the hash function can be avoided:

- A hash value can be computed once, stored, and used many times.
- Some hash functions can be computed more efficiently for a related set of strings. An example is the Karp–Rabin hash function.

We can further improve string binary search using precomputed information about the lcp’s between the strings in $R$.

Consider again the basic situation during string binary search:

- We want to compare $P$ and $S_{\text{mld}}$.
- We have already compared $P$ against $S_{\text{left}}$ and $S_{\text{right}}$, and we know $lcp(S_{\text{left}}, P)$ and $lcp(P, S_{\text{right}})$.

The values $left$ and $right$ are fully determined by $mid$ independently of $P$. That is, $P$ only determines whether the search ends up at position $mid$ at all, but if it does, $left$ and $right$ are always the same.

Thus, we can precompute and store the values $LLCP[\text{mid}] = lcp(S_{\text{left}}, S_{\text{mid}})$

$RLCP[\text{mid}] = lcp(S_{\text{mid}}, S_{\text{right}})$

Now we know all lcp values between $P$, $S_{\text{left}}$, $S_{\text{mid}}$, $S_{\text{right}}$ except $lcp(P, S_{\text{mid}})$. The following lemma shows how to utilize this.

**Algorithm 1.41:** String binary search (with precomputed lcps)

**Input:** Ordered string set $R = \{S_1, S_2, \ldots, S_n\}$, arrays $LLCP$ and $RLCP$, query string $P$.

**Output:** The number of strings in $R$ that are smaller than $P$.

1. $left \leftarrow 0$; $right \leftarrow n + 1$ (2) $lcp \leftarrow 0$; $rlcp \leftarrow 0$
3. While $right - left > 1$ do
   4. $mid \leftarrow ((left + right)/2)$ (5) $mclp \leftarrow min(lcp, rlcp)$
   6. $(x, mlncp) \leftarrow LcpCompare(P, S_{\text{mid}})$ (7) if $x = "<"$ then $right \leftarrow mid$; $rlcp \leftarrow mclp$
   8. Else $left \leftarrow mid$; $lcp \leftarrow mlp$
9. Return $left$

**Hashing and Fingerprints**

Hashing is a powerful technique for dealing with strings based on mapping each string to an integer. Hash tables come in many flavors that can be used with strings as well as with any other type of object with an appropriate hash function. A drawback of using a hash table to store a set of strings is that they do not support lcp and prefix queries. Hashing is also used in other situations, where one needs to check whether two strings $S$ and $T$ are the same or not:

- If $H(S) \neq H(T)$, then we must have $S \neq T$.
- If $H(S) = H(T)$, then $S = T$ and $S \neq T$ are both possible. If $S \neq T$, this is called a collision.

When used this way, the hash value is often called a fingerprint, and its range $[0, q)$ is typically large as it is not restricted by a hash table size.

**Definition 1.43:** The Karp–Rabin hash function for a string $S = s_0s_1 \ldots s_{m-1}$ over an integer alphabet is $H(S) = \sum_{i=0}^{m-1} s_i r^{m-1-i} \mod q$ for some fixed positive integers $q$ and $r$.

**Lemma 1.44:** For any two strings $A$ and $B$,

$H(AB) = (H(A) \cdot r^{|B|} + H(B)) \mod q$

$H(B) = (H(AB) - H(A) \cdot r^{|B|}) \mod q$

**Proof:** Without the modulo operation, the result would be obvious. The modulo does not interfere because of the rules of modular arithmetic:

$(x + y) \mod q = ((x \mod q) + (y \mod q)) \mod q$

$(xy) \mod q = ((x \mod q)(y \mod q)) \mod q$

Thus we can quickly compute $H(AB)$ from $H(A)$ and $H(B)$, and $H(B)$ from $H(AB)$ and $H(A)$. We will see applications of this later.

If $q$ and $r$ are coprime, then $r$ has a multiplicative inverse $r^{-1}$ modulo $q$, and we can also compute $H(A) = ((H(AB) - H(B)) \cdot (r^{-1})^{|B|}) \mod q$. 


The parameters \( q \) and \( r \) have to be chosen with some care to ensure that collisions are rare for any reasonable set of strings.

- The original choice is \( r = \sigma \) and \( q \) is a large prime.
- Another possibility is that \( q \) is a power of two and \( r \) is a small prime (\( r = 37 \) has been suggested). This is faster in practice, because the slow modulo operations can be replaced by bitshift operations. If \( q = 2^\sigma \), where \( \sigma \) is the machine word size, the modulo operations can be omitted completely.
- If \( q \) and \( r \) were both powers of two, then only the last \((\log_2 q) / \log_2 r\) characters of the string would affect the hash value. More generally, \( q \) and \( r \) should be coprime, i.e., have no common divisors other than 1.
- The hash function can be randomized by choosing \( q \) or \( r \) randomly. For example, if \( q \) is a prime and \( r \) is chosen uniformly at random from \([0, q)\), the probability that two strings of length \( m \) collide is at most \( m^2 / q \).
- A random choice over a set of possibilities has the additional advantage that we can change the choice if the first choice leads to too many collisions.

### Automata

Finite automata are a well known way of representing sets of strings. In this case, the set is often called a (regular) language.

- A trie is a special type of an automaton.
  - The root is the initial state, the leaves are accept states, ...
  - Trie is generally not a minimal automaton.
  - Trie techniques including path compaction can be applied to automata.
- Automata are much more powerful than tries in representing languages:
  - Infinite languages
  - Nondeterministic automata
  - Even an acyclic, deterministic automaton can represent a language of exponential size.
- Automata support set inclusion testing but not other trie operations:
  - No insertions and deletions
  - No satellite data, i.e., data associated to each string

### Sets of Strings: Summary

Efficient algorithms and data structures for sets of strings:

- Storing and searching: trie and ternary trie and their compact versions,
- Sorting: string quicksort and mergesort, LSD and MSD radix sort.

### 2. Exact String Matching

Let \( T = T[0..n) \) be the text and \( P = P[0..m) \) the pattern. We say that \( P \) occurs in \( T \) at position \( j \) if \( T[j..j + m) = P \). This is faster in practice, because the slow modulo operations can be replaced by bitshift operations. If \( q = 2^\sigma \), where \( \sigma \) is the machine word size, the modulo operations can be omitted completely.

- The worst case time complexity is \( O(nm) \). This happens, for example, when \( P = a^{m-1}b = aaaa..ab \) and \( T = a^m = aaaaaa..aa \).

The naive, brute force algorithm compares \( P \) against \( T[0..m) \), then against \( T[1..m) \), then against \( T[2..2 + m) \) etc. until an occurrence is found or the end of the text is reached. The text factor \( T[j..j + m) \) that is currently being compared against the pattern is called the text window.

#### Algorithm 2.2: Brute force

Input: \( T = T[0..n) \), pattern \( P = P[0..m) \)

Output: position of the first occurrence of \( P \) in \( T \)

1. \( i \leftarrow 0; j \leftarrow 0 \)
2. while \( i < m \) and \( j < n \)
3.   if \( P[j..j + m) = T[i..i + m) \) then \( i \leftarrow i + 1; j \leftarrow j + 1 \)
4.   else \( j \leftarrow j + 1 \)
5. if \( i = m \) then return \( j - m \) else return \( n \)

The worst case time complexity is \( O(nm) \). This happens, for example, when \( P = a^{m-1}b = aaaa..ab \) and \( T = a^m = aaaaaa..aa \).

### (Knuth–)Morris–Pratt

The Brute force algorithm forgets everything when it shifts the window. The Morris–Pratt (MP) algorithm remembers matches. It never goes back to a text character that already matched. The Knuth–Morris–Pratt (KMP) algorithm remembers mismatches too.

#### Example 2.3:

<table>
<thead>
<tr>
<th>Brute force</th>
<th>Morris–Pratt</th>
<th>Knuth–Morris–Pratt</th>
</tr>
</thead>
<tbody>
<tr>
<td>ainaisesti-ainainen</td>
<td>ainaisesti-ainainen</td>
<td>ainaisesti-ainainen</td>
</tr>
<tr>
<td>ainaiset</td>
<td>ainaiset</td>
<td>ainaiset</td>
</tr>
<tr>
<td>aina (6 comp.)</td>
<td>aina</td>
<td>aina</td>
</tr>
<tr>
<td>Aina (1)</td>
<td>Aina (1)</td>
<td>Aina (1)</td>
</tr>
<tr>
<td>Aina (1)</td>
<td>Aina (1)</td>
<td>Aina (1)</td>
</tr>
<tr>
<td>Aina (1)</td>
<td>Aina (1)</td>
<td>Aina (1)</td>
</tr>
</tbody>
</table>

We will describe the MP failure function here. The KMP failure function is left for the exercises.

- When the algorithm finds a mismatch between \( P[i] \) and \( T[j] \), we know that \( P[i..i + \delta = T[j..j - \delta) \). Specifically, we want to find the largest such \( \delta \).
- This means that \( P[i..i + \delta = T[j - \delta..j-j - \delta) \). In other words, \( P[i..i + \delta \) is the longest proper border of \( P[0..i] \).

#### Example: \( aib \) is the longest proper border of \( aina \).

- Thus \( fail[0] = 1 \) if \( P[0..0] = \epsilon \) has no proper border. We set \( fail[0] = -1 \).

- The only difference between MP and KMP is how they compute the failure function \( fail[i] \).

#### Algorithm 2.4: Knuth–Morris–Pratt / Morris–Pratt

Input: \( T = T[0..n) \), pattern \( P = P[0..m) \)

Output: position of the first occurrence of \( P \) in \( T \)

1. compute \( fail[0] \)
2. \( i \leftarrow 0; j \leftarrow 0 \)
3. while \( i < m \) and \( j < n \)
4.   if \( i < m \) and \( j < n \)
5.     if \( i = m \) then return \( j - m \) else return \( n \)
6.     if \( fail[i] = -1 \) means that there is no more pattern positions to compare against this text positions and we should move to the next text position.
7.     \( fail[i] = \) never needed here, but if we wanted to find all occurrences, it would tell how to continue after a full match.
Example 2.5: Let $P = \text{ainainen}$.

<table>
<thead>
<tr>
<th>$i$</th>
<th>$P[0..i)$</th>
<th>border</th>
<th>$\text{fail}[i]$</th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>$\varepsilon$</td>
<td>$-1$</td>
<td></td>
</tr>
<tr>
<td>1</td>
<td>$a$</td>
<td>$0$</td>
<td></td>
</tr>
<tr>
<td>2</td>
<td>$ai$</td>
<td>$0$</td>
<td></td>
</tr>
<tr>
<td>3</td>
<td>$ain$</td>
<td>$0$</td>
<td></td>
</tr>
<tr>
<td>4</td>
<td>$aina$</td>
<td>$0$</td>
<td></td>
</tr>
<tr>
<td>5</td>
<td>$ainai$</td>
<td>$3$</td>
<td></td>
</tr>
<tr>
<td>6</td>
<td>$ainai$</td>
<td>$0$</td>
<td></td>
</tr>
<tr>
<td>7</td>
<td>$ainaine$</td>
<td>$0$</td>
<td></td>
</tr>
</tbody>
</table>

The $(K)MP$ algorithm operates like an automaton, since it never moves backwards in the text. Indeed, it can be described by an automaton that has a special failure transition, which is an $\varepsilon$-transition that can be taken only when there is no other transition to take.

Theorem 2.7: Algorithms MP and KMP preprocess a pattern in time $O(m)$ and then search the text in time $O(n)$ for ordered alphabet.

Proof. We show that the text search requires $O(n)$ time. Exactly the same argument shows that pattern preprocessing needs $O(m)$ time.

It is sufficient to count the number of comparisons that the algorithms make. After each comparison $P[i]$ vs. $T[j]$, one of the two conditional branches is executed:

1. Here $j$ is incremented. Since $j$ never decreases, this branch can be taken at most $n + 1$ times.
2. Here $i$ decreases since $\text{fail}[i] < i$. Since $i$ only increases in the then-branch, this branch cannot be taken more often than the then-branch.

Let $w$ be the wordsize of the computer, typically 64. Assume first that $m \leq w$. Then each bitvector can be stored in a single integer and the bitwise operations can be executed in constant time.

Algorithm 2.8: Shift-And

Input: text $T = T[0 \ldots n]$, pattern $P = P[0 \ldots m]$.

Output: position of the first occurrence of $P$ in $T$.

Preprocess:

1. For $c \in \Sigma$ do $B[c] \leftarrow 0$.
2. For $i \leftarrow 0$ to $m - 1$ do $B[P[i]] \leftarrow B[P[i]] + 2^i$ // $B[P[i]] = i + 1$.

Search:

1. $D \leftarrow 0$.
2. $j \leftarrow 0$.
3. If $D < m$ do $D \leftarrow (D + 1) \& B[T[j]]$.
4. If $D = 0$ do $j \leftarrow j + 1$.
5. $D \leftarrow (D < m) \& B[T[j]]$.
6. If $D = 0$ then return $j - m + 1$ // $D \leftarrow (D - 1)$.
7. Return $n$.

Shift-Or is a minor optimization of Shift-And. It is the same algorithm except the roles of 0’s and 1’s in the bitvectors have been swapped. Then line 5 becomes $D \leftarrow (D < m) \& B[T[j]]$. Note that the “|1” was removed, because the shift already brings the correct bit to the least significant bit position.

On an integer alphabet when $m \leq w$:

- Preprocessing time is $O(n)$.
- Search time is $O(n)$.

If $m > w$, we can store the bitvectors in $[m/w]$ machine words and perform each bitvector operation in $O(m/w)$ time.

- Preprocessing time is $O(m/w + m)$.
- Search time is $O((m/w) + m)$.

If no pattern prefix longer than $w$ matches a current text suffix, then only the least significant machine word contains 1’s. There is no need to update the other words; they will stay 0.

- Then the search time is $O(n)$ on average.

Algorithms like Shift-And that take advantage of the implicit parallelism in bitvector operations are called bitparallel.

An efficient algorithm for computing the failure function is very similar to the search algorithm itself:

- In the MP algorithm, when we find a match $P[i] = T[j]$, we know that $P[0..i] = T[j - i]$. More specifically, $P[0..i]$ is the longest prefix of $P$ that matches a suffix of $T[0..j]$.
- Suppose $T = \#P[m..n]$, where $\#$ is a symbol that does not occur in $P$. Finding a match $P[i] = T[j]$, we know that $P[0..j]$ is the longest prefix of $P$ that is a proper suffix of $P[0..j]$. Thus $\text{fail}[j + 1] = i + 1$.

Algorithm 2.6: Morris-Pratt failure function computation

Input: pattern $P = P[0 \ldots m]$.

Output: array $\text{fail}[m]$ for $P$.

1. $i \leftarrow -1$; $\text{fail}[i] \leftarrow i$.
2. While $j < m$ do
3. If $j = -1$ or $P[i] = P[j]$ then $i \leftarrow i + 1$; $j \leftarrow j + 1$; $\text{fail}[j] \leftarrow i$.
4. Else $i \leftarrow \text{fail}[i]$.
5. Return $\text{fail}$.

When the algorithm reads $\text{fail}[i]$ on line 4, $\text{fail}[i]$ has already been computed.

Shift-And (Shift-Or)

When the MP algorithm is at position $j$ in the text $T$, it computes the longest prefix of the pattern $P[0..m]$ that is a suffix of $T[0..j]$. The Shift-And algorithm computes all prefixes of $P$ that are suffixes of $T[0..j]$.

- The information is stored in a bitvector $D$ of length $m$, where $D_i = 1$ if $P[0..i] = T[j - i]$ and $D_i = 0$ otherwise. $(D_0$ is the least significant bit $)$.
- When $D.(m - 1) = 1$, we have found an occurrence.

The bitvector $D$ is updated at each text position $j$:

- There are precomputed bitvectors $B[c]$, for all $c \in \Sigma$, where $B[c].i = 1$ if $P[i] = c$ and $B[c].i = 0$ otherwise.
- $D$ is updated in two steps:
  1. $D \leftarrow (D \ll 1)$ (the bitwise shift and the bitwise or). Now $D$ tells, which bitvectors would match if $T[j]$ would match every character.
  2. $D \leftarrow D \& B[T[j]]$ (the bitwise and). Remove the prefixes where $T[j]$ does not match.

Example 2.9: $P = \text{assai}$, $T = \text{apassi}$, bitvectors are columns.

$B[c]$, $c \in \{a, i, p, s\}$

<table>
<thead>
<tr>
<th>$B[c]$</th>
<th>$D$ at each step</th>
<th>$D$ at each step</th>
</tr>
</thead>
<tbody>
<tr>
<td></td>
<td>$a$</td>
<td>$1$</td>
</tr>
<tr>
<td>$a$</td>
<td>$1$</td>
<td>$0$</td>
</tr>
<tr>
<td>$i$</td>
<td>$0$</td>
<td>$0$</td>
</tr>
<tr>
<td>$p$</td>
<td>$0$</td>
<td>$0$</td>
</tr>
<tr>
<td>$s$</td>
<td>$0$</td>
<td>$0$</td>
</tr>
</tbody>
</table>

The Shift-And algorithm can also be seen as a bitparallel simulation of the nondeterministic automaton that accepts a string ending with $P$.

Karp–Rabin

The Karp–Rabin hash function (Definition 1.43) was originally developed for solving the exact string matching problem. The idea is to compute the hash values or fingerprints $H(P)$ and $H(T[j..j + m])$ for all $j \in [0..n - m]$.

- If $H(P) \neq H(T[j..j + m])$, then we must have $P \neq T[j..j + m]$.
- If $H(P) = H(T[j..j + m])$, the algorithm compares $P$ and $T[j..j + m]$ in brute force manner. If $P \neq T[j..j + m]$, this is a false positive.

The text factor fingerprints are computed in a sliding window fashion. The fingerprint for $T[j + 1..j + 1 + m]$ is $a(T[j..j + m])$ is computed from the fingerprint for $T[j..j + m]$ and $T[j + 1..j + 1 + m]$. If $P \neq T[j..j + m]$, this is a false positive.

A hash function that supports this kind of sliding window computation is known as a rolling hash function.
Algorithm 2.10: Karp-Rabin
Input: text $T = T[0...n)$, pattern $P = P[0...m)$
Output: position of the first occurrence of $P$ in $T$

1. Choose $q$ and $r$; $s ← r^{m−1} \mod q$
2. $hp ← 0$, $ht ← 0$
3. for $i ← 0$ to $m − 1$ do $hp ← (hp \cdot r + P[i]) \mod q$ // $hp = H(P)$
4. for $j ← 0$ to $m − 1$ do $ht ← (ht \cdot r + T[j]) \mod q$
5. for $j ← 0$ to $n − m − 1$ do
6. if $hp = ht$ then if $P = T[j...j+m)$ then return $j$
7. $ht ← (ht − (T[j−1] \cdot s) \cdot r + T[j+m]) \mod q$
8. if $hp = ht$ then if $P = T[j...j+m)$ then return $j$
9. return $n$

On an integer alphabet:
- The worst case time complexity is $O(mn)$.
- The average case time complexity is $O(m + n)$.

Karp–Rabin is not competitive in practice for a single pattern, but can be for multiple patterns (exercise).

More precisely, suppose we are currently comparing $P$ against $T[j..j+m)$. Start by comparing $P[m−1]$ to $T[k]$, where $k = j + m − 1$.

- If $P[m−1] \neq T[k]$, shift the pattern until the pattern character aligned with $T[k]$ matches, or until the full pattern is past $T[k]$.
- If $P[m−1] = T[k]$, compare the rest in a brute force manner. Then shift to the next position, where $T[k]$ matches.

The length of the shift is determined by the shift table that is precomputed for the pattern. $shift[c]$ is defined for all $c \in \Sigma$:
- If $c$ does not occur in $P$, $shift[c] = m$.
- Otherwise, $shift[c] = m − 1 − i$, where $P[i] = c$ is the last occurrence of $c$ in $P[0..m−2]$.

Example 2.12: $P = \text{ainainen}$.

<table>
<thead>
<tr>
<th>$c$</th>
<th>last occ.</th>
<th>$shift$</th>
</tr>
</thead>
<tbody>
<tr>
<td>a</td>
<td>ainainen</td>
<td>4</td>
</tr>
<tr>
<td>i</td>
<td>ainainen</td>
<td>1</td>
</tr>
<tr>
<td>n</td>
<td>ainainen</td>
<td>3</td>
</tr>
<tr>
<td>e</td>
<td>ainainen</td>
<td>2</td>
</tr>
<tr>
<td>$\Sigma \setminus {a,e,i,n}$</td>
<td>-</td>
<td>8</td>
</tr>
</tbody>
</table>

On an integer alphabet:
- Preprocessing time is $O(\sigma + m)$.
- In the worst case, the search time is $O(mn)$.
  - For example, $P = b^a$ and $T = a^a$.
- In the best case, the search time is $O(n/m)$.
  - For example, $P = b^n$ and $T = a^n$.
- In the average case, the search time is $O(\sigma \cdot m/n \cdot n)$.
  - This assumes that each pattern and text character is picked independently by uniform distribution.

In practice, a tuned implementation of Horspool is very fast when the alphabet is not too small.

Factor based algorithms use an automaton that accepts suffixes of the reverse pattern $P^R$ (or equivalently reverse prefixes of the pattern $P$).

- BDM (Backward DAWG Matching) uses a deterministic automaton that accepts exactly the suffixes of $P^R$.
- DA WG (Directed Acyclic Word Graph) is also known as suffix automaton.
- BN DM (Backward Nondeterministic DAWG Matching) simulates a nondeterministic automaton.

Example 2.15: $P = \text{assin}$.

![Diagram]

- BOM (Backward Oracle Matching) uses a much simpler deterministic automaton that accepts all suffixes of $P^R$ but may also accept some other strings. This can cause shorter shifts but incorrect behaviour.

Horspool

The algorithms we have seen so far access every character of the text. If we start the comparison between the pattern and the current text position from the end, we can often skip some text characters completely.

There are many algorithms that start from the end. The simplest are the Horspool-type algorithms.

The Horspool algorithm checks first the last character of the text window, i.e. the character aligned with the last pattern character. If that doesn’t match, it moves (shifts) the pattern forward until there is a match.

Example 2.11:

<table>
<thead>
<tr>
<th>Pattern</th>
<th>Shift</th>
</tr>
</thead>
<tbody>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
</tbody>
</table>

Algorithm 2.13: Horspool
Input: text $T = T[0...n)$, pattern $P = P[0...m)$
Output: position of the first occurrence of $P$ in $T$

Preprocess:
1. for $c \in \Sigma$ do $shift[c] ← m$
2. for $i ← 0$ to $m − 1$ do $shift[P[i]] ← m − 1 − i$

Search:
3. $j ← 0$
4. while $j + m \leq n$ do
5. if $P[m−1] = T[j+m−1]$ then
6. $i ← m − 2$
7. while $i > 0$ and $P[i] = T[j+i]$ do $i ← i − 1$
8. if $i = 1$ then return $j$
9. $j ← j + shift[T[j+m−1]]$
10. return $n$

BN DM

Starting the matching from the end enables long shifts.

- The Horspool algorithm bases the shift on a single character.
- The Boyer–Moore algorithm uses the matching suffix and the mismatching character.
- Factor based algorithms continue matching until no pattern factor matches. This may require more comparisons but it enables longer shifts.

Example 2.14:

<table>
<thead>
<tr>
<th>Pattern</th>
<th>Shift</th>
</tr>
</thead>
<tbody>
<tr>
<td>varasti-ailainen</td>
<td>varasti-ailainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
<tr>
<td>ainainen</td>
<td>ainainen</td>
</tr>
</tbody>
</table>

Suppose we are currently comparing $P$ against $T[j...j+m)$. We use the automaton to scan the text backwards from $T[j+m−1]$. When the automaton has scanned $T[j+i..j+m)$:

- If the automaton is in an accept state, then $T[j+i..j+m)$ is a prefix of $P$.
  → If $i = 0$, we found an occurrence.
  → Otherwise, mark the prefix match by setting $shift = i$. This is the length of the shift that would achieve a matching alignment.
- If the automaton can still reach an accept state, then $T[j+i..j+m)$ is a factor of $P$.
  → Continue scanning.
- When the automaton can no more reach an accept state:
  → Stop scanning and shift: $j ← j + shift$. 


BNDM does a bitparallel simulation of the nondeterministic automaton, which is quite similar to Shift-And.

The state of the automaton is stored in a bitvector $D$. When the automaton has scanned $T[j + i j + m]$, we have:
- $D k = 1$ if and only if there is a path from the initial state to state $k$ with the string $(T[j + i j + m])^k$, and thus $T[j + i j + m] = P[m - k - 1.2m - k - i - 1]$.
- If $D(m - 1) = 1$, then $T[j + i j + m]$ is a prefix of the pattern.
- If $D = 0$, then the automaton can no more reach an accept state.

Updating $D$ uses precomputed bitvectors $B[c]$, for all $c \in \Sigma$:
- $B[c] i = 1$ if and only if $P[m - 1 - i] = T^R[i] = c$.

The update when reading $T[j + i]$ is familiar: $D = (D < < 1) \& B[T[j + i]]$.

- Note that there is no "1". This is because $D(1) = 0$ always after reading at least one character, so the shift brings the right bit to $D.0$.
- Before reading anything $D.(-1) = 1$. This exception is handled by starting the computation with the first shift already performed. Because of this, the shift is done at the end of the loop.

Example 2.17: Let $S[0..m)$ be a string. An integer $p \in [1..m)$ is a period of $S$ if $S[i] = S[i + p]$ for all $i \in [0..m - p)$. The smallest period of $S$ is denoted $\text{per}(S)$. $S$ is $k$-periodic if $m \geq k \cdot \text{per}(S)$.

Definition 2.18: Let $S[0..m)$ be a string. An integer $p \in [1..m]$ is a period of $S$, if $S[i] = S[i + p]$ for all $i \in [0..m - p)$. The smallest period of $S$ is denoted $\text{per}(S)$. $S$ is $k$-periodic if $m \geq k \cdot \text{per}(S)$.

Example 2.19: The periods of $S_1 = \text{asabaabas}$ are 4, 7, 8 and 9. The periods of $S_2 = \text{ababcababcaba}$ are 3, 6, 9, 12 and 13. $S_2$ is 3-periodic but $S_1$ is not.

There is a strong connection between periods and borders.

Lemma 2.20: $p$ is a period of $S[0..m)$ if and only if $S$ has a proper border of length $m - p$.

Proof. Both conditions hold if and only if $S[0..m - p) = S[p..m)$. □

Corollary 2.21: The length of the longest proper border of $S$ is $m - \text{per}(S)$.

Recall that $\text{fail}[i]$ in MP is the length of the longest proper border of $P[0..i)$. Thus the pattern shift by MP is $\ell - \text{fail}[\ell] = \text{per}(P[0..\ell))$ and the lcp skip is $\text{fail}[i] = \ell - \text{per}(P[0..\ell))$. Thus knowing $\text{per}(P[0..\ell))$ is sufficient to emulate MP shift and skip.

The Crochemore algorithm has two cases:
- If $P[0..\ell)$ is 3-periodic, then compute $\text{per}(P[0..\ell))$ and do the MP shift and skip.
- If $P[0..\ell)$ is not 3-periodic, then shift by $(\ell/3) + 1 \leq \text{per}(P[0..\ell))$ and start the lcp comparison from scratch.

To find out if $P[0..\ell)$ is 3-periodic and to compute $\text{per}(P[0..\ell))$ if it is, Crochemore uses another combinatorial concept.

### Algorithm 2.16: BNDM

**Input:** text $T \equiv T[0..n)$, pattern $P \equiv P[0..m)$

**Output:** position of the first occurrence of $P$ in $T$

**Preprocess:**
1. for $c \in \Sigma$ do $B[c] \leftarrow 0$
2. for $i \leftarrow 0$ to $m - 1$ do $B[P[m - 1 - i]] \leftarrow B[P[m - 1 - i]] + 2^i$

**Search:**
3. $j \leftarrow 0$
4. while $j + m \leq n$ do
5. $i \leftarrow m$; $\text{shift} = m$
6. $D \leftarrow 2^m - 1$ // $D \leftarrow 1$
7. while $D \neq 0$ do
8. // Now $T[j + i j + m]$ is a pattern factor
9. $i \leftarrow i - 1$
10. $D \leftarrow D \& B[T[j + i]]$
11. if $i = 0$ then return $j$
12. else $\text{shift} = i$
13. $D \leftarrow D < < 1$
14. $j \leftarrow \text{shift} + j$
15. return $n$
**Definition 2.22:** Let $MS(S)$ denote the lexicographically maximal suffix of a string $S$. If $S = MS(S)$, $S$ is called **self-maximal**.

Period computation is easier for maximal suffixes and self-maximal strings than for arbitrary strings.

**Lemma 2.23:** Let $S[0..m]$ be a self-maximal string and let $p = per(S)$. For any $r \in \Sigma$,

- $MS(S\{c\}) = S[c..m]$ and $per(S\{c\}) = p$ if $c = S[m-p]$.
- $MS(S\{c\}) = S[c..m]$ and $per(S\{c\}) = m+1$ if $c < S[m-p]$.
- $MS(S\{c\}) \neq S[c..m]$ if $c > S[m-p]$.

Furthermore, let $r = m \mod p$ and $R = S[m-r..m)$. Then $R$ is self-maximal and $MS(S\{c\}) = MS(R\{c\})$ if $c > S[m-p]$.

As the final piece of the Crochemore algorithm, the following result shows how to use the maximal suffix information to obtain information about the periodicity of the full string.

**Lemma 2.25:** Let $S[0..m]$ be a string and let $S[0..m] = MS(S)$ and $p = per(MS(S))$.

- $S$ is 3-periodic if and only if $p \leq m/3$ and $S[0..p) = S[p..p+s)$.
- If $S$ is 3-periodic, then $per(S) = p$.

The algorithm is given on the next slide.

For ordered alphabet:

- The time complexity is $O(n)$.
- The algorithm uses only a constant number of integer variables in addition to the strings $P$ and $T$.

Crochemore is not competitive in practice. However, there are situations, where the pattern can be very long and the space complexity is more important than speed.

There are also other linear time, constant extra space algorithms. All of them are based on string periodicity in some way.

**Algorithm 2.24:** Update-$MS(P, l, s, p)$

**Input:** a string $P$ and integers $l, s, p$ such that $MS(P[0..l]) = P[l..s)$ and $p = per(P[s..l])$.

**Output:** a tuple $(l'+1, s'+1)$ such that $MS(P[0..l'+1]) = P[l'+1..s'+1)$ and $p' = per(P[s'+1..l'+1))$.

1. if $l = 0$ then return $(1, 0, 1)$
2. $i \leftarrow l$
3. while $i < l + 1$ do
   4. if $P[i] > P[l - p]$ then
      5. $i \leftarrow i + 1$
   6. else if $P[i] < P[l - p]$ then
      7. $i \leftarrow i - 1$
   8. else $i \leftarrow i + 1$
4. return $(l + 1, s, p)$

**Algorithm 2.26:** Crochemore

**Input:** strings $T[0..n]$ (text) and $P[0..m]$ (pattern).

**Output:** position of the first occurrence of $P$ in $T$

1. $j \leftarrow l \leftarrow p \leftarrow s \leftarrow 0$
2. while $j + m \leq n$ do
   3. while $j < l$ and $l < m$ and $T[j..j + l] = P[l..l + m]$ do
      4. $(t, s, p) \leftarrow Update-MS(P, l, s, p)$
   5. if $l = m$ then return $j$
   6. if $p < l/3$ and $P[p..p + s) = P[0..p + s)$ then
      7. $j \leftarrow j + l$
   8. else if $P[0..l] > P[l..l + m] + 1$ then
      9. $j \leftarrow j + l/3 + 1$
   10. $(t, s, p) \leftarrow (0, 0, 0)$
11. return $n$

**Aho–Corasick**

Given a text $T$ and a set $P = \{P_1, P_2, \ldots, P_k\}$ of patterns, the **multiple exact string matching** problem asks for the occurrences of all the patterns in the text. The Aho–Corasick algorithm is an extension of the Morris–Pratt algorithm for multiple exact string matching.

Aho–Corasick uses the trie $trie(P)$ as an automaton and augments it with a failure function similar to the Morris–Pratt failure function.

**Example 2.27:** Aho–Corasick automaton for $P = \{\text{ba}, \text{she}, \text{his}, \text{hers}\}$.

At each stage of the matching, the algorithm computes the node $v$ such that $S_v$ is the longest suffix of $T[0..j]$ represented by any node.

**Algorithm 2.29:** Aho–Corasick

**Input:** text $T$, pattern set $P = \{P_1, P_2, \ldots, P_k\}$.

**Output:** all pairs $(i, j)$ such that $P_i$ occurs in $T$ ending at $j$.

1. $(root, child, fail, patterns) \leftarrow Construct-AC-Automaton(P)$
2. $v \leftarrow root$
3. for $j \leftarrow 0$ to $n - 1$ do
   4. while $child(v, T[j]) \neq \bot$ do $v \leftarrow fail(v)$
   5. $v \leftarrow child(v, T[j])$
   6. for $i \in patterns(v)$ do output $(i, j)$

The construction of the automaton is done in two phases: the trie construction and the failure links computation.

**Algorithm 2.30:** Construct-AC-Automaton

**Input:** pattern set $P = \{P_1, P_2, \ldots, P_k\}$. 

**Output:** AC automaton $(root, child, fail, patterns)$.

1. $(root, child, patterns) \leftarrow Construct-AC-Trie(P)$
2. $(fail, patterns) \leftarrow Compute-AC-Fail(root, child, patterns)$
3. return $(root, child, fail, patterns)$
Algorithm 2.31: Construct-AC-Trie
Input: pattern set \( P = \{P_1, P_2, \ldots, P_k\} \).
Output: AC trie: root, child() and patterns().

1. Create new node root
2. for \( i \leftarrow 1 \) to \( k \) do
3. \( \text{v} \leftarrow \text{root}, j \leftarrow 0 \)
4. while \( \text{child}(\text{v}, P[j]) \neq \perp \) do
5. \( \text{v} \leftarrow \text{child}(\text{v}, P[j]), j \leftarrow j + 1 \)
6. while \( j < |P_i| \) do
7. Create new node \( \text{u} \)
8. \( \text{child}(\text{v}, P[j]) \leftarrow \text{u} \)
9. \( \text{v} \leftarrow \text{w}, j \leftarrow j + 1 \)
10. patterns(\( \text{v} \)) = \{i|P_i \text{ is a suffix of } S_\text{v}\}
11. return (root, child(), patterns())

Lines (3)–(10) perform the standard trie insertion (Algorithm 1.2).

- Line (10) marks \( v \) as a representative of \( P_i \).
- The creation of a new node \( v \) initializes patterns(\( v \)) to \( \emptyset \) (in addition to initializing child(\( v, c \)) to \( \perp \) for all \( c \in \Sigma \)).

Algorithm 2.32: Compute-AC-Fail
Input: AC trie: root, child() and patterns()
Output: AC failure function fail() and updated patterns()

1. Create new node failback
2. for \( c \in \Sigma \) do child(failback, \( c \)) \leftarrow root
3. fail(root) \leftarrow failback
4. queue \leftarrow root
5. while queue \neq \emptyset do
6. \( u \leftarrow \text{popfront(queue)} \)
7. for \( c \in \Sigma \) such that \( \text{child}(u, c) \neq \perp \) do
8. \( v \leftarrow \text{child}(u, c) \)
9. \( w \leftarrow \text{fail}(u) \)
10. while \( \text{child}(w, c) \neq \perp \) do \( w \leftarrow \text{fail}(w) \)
11. fail(v) \leftarrow child(w, c)
12. patterns(v) \leftarrow patterns(v) \cup patterns(fail(v))
13. pushback(queue, v)
14. return (fail(), patterns())

The algorithm does a breath first traversal of the trie. This ensures that correct values of fail() and patterns() are already computed when needed.

Summary: Exact String Matching

Exact string matching is a fundamental problem in stringology. We have seen several different algorithms for solving the problem.

The properties of the algorithms vary with respect to worst case time complexity, average case time complexity, type of alphabet (ordered/integer) and even space complexity.

The algorithms use a wide range of completely different techniques:

- There exists numerous algorithms for exact string matching but most of them use variations or combinations of the techniques we have seen (study groups).
- Many of the techniques can be adapted to other problems. All of the techniques have some uses in practice.

Edit distance

The edit distance ed(\( A, B \)) of two strings \( A \) and \( B \) is the minimum number of edit operations needed to change \( A \) into \( B \). The allowed edit operations are:

- **S** Substitution of a single character with another character.
- **I** Insertion of a single character.
- **D** Deletion of a single character.

Example 3.1: Let \( A = \text{Levenshtein} \) and \( B = \text{Levenshtein} \). Then ed(\( A, B \)) = 3.

The set of edit operations can be described with an edit sequence:

<table>
<thead>
<tr>
<th>N</th>
<th>K</th>
<th>N</th>
<th>N</th>
<th>L</th>
<th>I</th>
<th>E</th>
</tr>
</thead>
</table>

or with an alignment:

- Levenshtein
- Levenshtein

In the edit sequence, \( N \) means No edit.

There are many variations and extension of the edit distance, for example:

- **Hamming distance** allows only the substitution operation.
- **Damerau-Levenshtein distance** adds an edit operation:

- **T** Transposition swaps two adjacent characters.

With **weighted edit distance**, each operation has a cost or weight, which can be other than one. **Non insertions and deletions (inds) of factors** at a cost that is lower than the sum of character indels.

We will focus on the basic Levenshtein distance.

Levenshtein distance has the following two useful properties, which are not shared by all variations (exercise):

- **Levenshtein distance is a metric**.
- If ed(\( A, B \)) = 3, there exists an edit sequence and an alignment with 3 edit operations, but no edit sequence or alignment with less than 3 edit operations. An edit sequence and an alignment with ed(\( A, B \)) edit operations is called **optimal**.

3. Approximate String Matching

Often in applications we want to search a text for something that is similar to the pattern but not necessarily exactly the same.

To formalize this problem, we have to specify what does “similar” mean. This can be done by defining a similarity or a distance measure.

A natural and popular distance measure for strings is the edit distance, also known as the Levenshtein distance.

Assuming \( \sigma \) is constant:

- The search time is \( O(n) \).
- The space complexity is \( O(m) \), where \( m = ||P|| \).
  - The implementation of patterns() requires care (exercise).
  - The preprocessing time is \( O(m) \), where \( m = ||P|| \).
  - The only non-trivial issue is the while-loop on line (10).

With \( \text{root}, v_1, v_2, \ldots, v_m \) be the nodes on the path from root to a node representing a pattern \( P_i \). Let \( w_j = \text{fail}(v_j) \) for all \( j \). Let depth(\( v \)) be the depth of a node \( v \) (depth(root) = 0).

When processing \( v_j \) and computing \( w_j = \text{fail}(v_j) \), we have depth(\( w_j \)) = depth(\( w_{j-1} \)) + 1 before line (10) and depth(\( w_j \)) < depth(\( w_{j-1} \)) + 1 after line (10), where \( j \) is the number of rounds in the while-loop.

Thus, the total number of rounds in the while-loop when processing the nodes \( v_1, v_2, \ldots, v_m \) is at most \( \ell = ||P|| \), and thus over the whole algorithm at most \( ||P|| \).

The analysis when \( \sigma \) is not constant is left as an exercise.
Computing Edit Distance

Given two strings $A[1..m]$ and $B[1..n]$, define the values $d_{ij}$ with the recurrence:

\[
d_{00} = \ldots \quad A[i] = B[j]
\]

Theorem 3.2: $d_{ij} = \text{ed}(A[1..i], B[1..j])$ for all $0 \leq i \leq m$, $0 \leq j \leq n$.

In particular, $d_{mn} = \text{ed}(A, B)$.

Example 3.3: $A=$ ballad, $B=$ handball

\[
\begin{array}{cccccccc}
 & b & a & n & d & b & a & l & l \\
0 & 1 & 1 & 2 & 3 & 4 & 5 & 6 & 7 \\
1 & 2 & 2 & 1 & 2 & 3 & 4 & 4 & 5 & 6 \\
2 & 1 & 3 & 3 & 2 & 2 & 3 & 4 & 4 & 5 & 6 \\
3 & 1 & 4 & 4 & 3 & 3 & 4 & 4 & 5 & 5 & 4 \\
4 & 1 & 5 & 5 & 4 & 4 & 4 & 4 & 5 & 5 & 5 \\
5 & 1 & 6 & 6 & 5 & 5 & 4 & 4 & 5 & 5 & 5 & 6 \\
\end{array}
\]

$\text{ed}(A, B) = d_{mn} = d_{6,8} = 6$.

Example 3.6: $A =$ ballad, $B =$ handball

\[
\begin{array}{cccccccc}
 & b & a & n & d & b & a & l & l \\
0 & 1 & 1 & 2 & 3 & 4 & 5 & 6 & 7 \\
1 & 2 & 2 & 1 & 2 & 3 & 4 & 4 & 5 & 6 \\
2 & 1 & 3 & 3 & 2 & 2 & 3 & 4 & 4 & 5 & 6 \\
3 & 1 & 4 & 4 & 3 & 3 & 4 & 4 & 5 & 5 & 4 \\
4 & 1 & 5 & 5 & 4 & 4 & 4 & 4 & 5 & 5 & 5 & 6 \\
\end{array}
\]

There are 7 paths from $(0,0)$ to $(6,8)$ corresponding to 7 different optimal edit sequences and alignments, including the following three:

IIIINNNNDD SNISSNIS SSNSIS

---ballad ba-lia-d ball-ad- handball-- handball handball
Approximate String Matching

Now we are ready to tackle the main problem of this part: approximate string matching.

Problem 3.7: Given a text \(T[1..n]\), a pattern \(P[1..m]\) and an integer \(k \geq 0\), report all positions \(j \in \{1..n\}\) such that \(ed(P,T(j–\ell..j)) \leq k\) for some \(\ell \geq 0\).

The factor \(T(j–\ell..j)\) is called an approximate occurrence of \(P\).

There can be multiple occurrences of different lengths ending at the same position \(j\), but usually it is enough to report just the end positions. We ask for the end position rather than the start position because that is more natural for the algorithms.

Proof. We use induction with respect to \(i+j\).

Basis: \(g_{00} = 0 = ed(\epsilon,\epsilon) = ed(\epsilon,T(j–0..j))\) (min at \(\ell = 0\))

\(g_{0j} = 0 = ed(\epsilon,\epsilon) = ed(\epsilon,T(0–0..0))\) (\(0 \leq \ell \leq j = 0\))

Induction step: Essentially the same as in the proof of Theorem 3.2.

Algorithm 3.10: Approximate string matching

Input: text \(T[1..n]\), pattern \(P[1..m]\), and integer \(k\)

Output: end positions of all approximate occurrences of \(P\)

1. for \(i \leftarrow 0\) to \(m\) do \(g_{0i} \leftarrow i\)
2. for \(j \leftarrow 1\) to \(n\) do \(g_{0j} \leftarrow 0\)
3. for \(j \leftarrow 1\) to \(n\) do
4. for \(i \leftarrow 1\) to \(m\) do
5. \(g_{ij} \leftarrow \min\{g_{i–1,j–1} + \delta(A[i],B[j]),g_{i–1,j} + 1,g_{i,j–1} + 1\}\)
6. if \(g_{mj} \leq k\) then output \(j\).

* Time and space complexity is \(O(mn)\) on ordered alphabet.
* The space complexity can be reduced to \(O(n)\) by storing only one column as in Algorithm 3.5.

Lemma 3.12: For every \(i \in \{1..m\}\) and every \(j \in \{1..n\}\),

\(g_i = g_{i–1,j–1} + 1\) or \(g_i = g_{i–1,j} + 1\).

Proof. By definition, \(g_{ij} \leq g_{i–1,j–1} + \delta(P[i],T[j]) \leq g_{i–1,j–1} + 1\). We show that \(g_{ij} \geq g_{i–1,j–1}\) by induction on \(i+j\).

The induction assumption is that \(g_{ij} \geq g_{i–1,j–1}\) when \(p \in \{1..m\}, q \in \{1..n\}\) and \(p + q < i+j\). At least one of the following holds:

1. \(g_i = g_{i–1,j–1} + \delta(P[i],T[j])\). Then \(g_{ij} \geq g_{i–1,j–1}\).
2. \(g_i = g_{i–1,j} + 1\) and \(i+1\). Then \(g_{ij} = g_{i–1,j} + 1 \geq g_{i–2,j–1} + 1 \geq g_{i–1,j–1}\).
3. \(g_i = g_{i–1,j} + 1\) and \(j > 1\). Then \(g_{ij} = g_{i–1,j} + 1 \geq g_{i–2,j–1} + 1 \geq g_{i–1,j–1}\).
4. \(g_i = g_{i–1,j} + 1\) and \(i=1\). Then \(g_{ij} = 0 + 1 = g_{i–1,j–1}\).
5. \(g_i = g_{i–1,j} + 1\) and \(j=1\). Then \(g_{ij} = i+1 = (i–1) + 2 = g_{i–1,j–1} + 2\),

which cannot be true. Thus this case can never happen. □

Define the values \(g_j\) with the recurrence:

\(g_0 = 0, 0 \leq j \leq n\)
\(g_i = i, 1 \leq i \leq m\)
\(g_{ij} = \min\{g_{i–1,j–1} + \delta(P[i],T[j]),g_{i–1,j} + 1,g_{i,j–1} + 1\}\)

for \(1 \leq i \leq m, 1 \leq j \leq n\).

Theorem 3.8: For all \(0 \leq i \leq m, 0 \leq j \leq n\):

\(g_{ij} = \min\{ed(P[1..i],T(j–\ell..j)) | 0 \leq \ell \leq j\}.

In particular, \(j\) is an ending position of an approximate occurrence if and only if \(g_{mj} \leq k\).

Example 3.9: \(P = \text{match}, T = \text{remachine}, k = 1\)

```
g r e m a c h i n e
0 0 0 0 0 0 0 0 0 0
m ⇒
1 1 1 0 1 1 1 1 1 1
a ⇒
2 2 2 1 0 1 2 2 2 2
m t ⇒
3 3 3 2 1 1 2 3 3 3
a c ⇒
4 4 4 3 2 1 2 3 4 4
h ⇒
5 5 5 4 3 2 1 2 3 4
```

One occurrence ending at position 6.

Ukkonen’s Cut-off Heuristic

We can speed up the algorithm using the diagonal monotonicity of the matrix \((g_{ij})\):

A diagonal \(d, –m \leq d \leq n\), consists of the cells \(g_{ij}\) with \(j–i = d\).

Every diagonal \((g_{ij})\) is monotonically non-decreasing.

Example 3.11: Diagonals -3 and 2.

```
g a t a s t r u c t u r e
0 0 0 0 0 0 0 0 0 0
m a t c h
1 1 1 1 0 1 1 1 1 1
2 2 2 1 0 1 2 2 2 2
m t c h
3 3 3 2 1 1 2 3 3 3
a t c h
4 4 4 3 2 1 2 3 4 4
h t c h
5 5 5 4 3 2 1 2 3 4
```

We can reduce computation using diagonal monotonicity:

* Whenever the value on a diagonal \(d\) grows larger than \(k\), we can discard \(d\) from consideration, because we are only interested in values at most \(k\) on the row \(m\).
* We keep track of the smallest undiscarded diagonal \(d\). Each column is computed only up to diagonal \(d+1\).

Example 3.13: \(P = \text{strict}, T = \text{datastructure}, k = 1\)

```
g d a t a s t r u c t u r e
0 0 0 0 0 0 0 0 0 0
s t r i c t
1 1 1 1 1 0 1 1 1 1 1
2 2 2 1 2 1 0 1 2 2 2
m t r i c t
2 2 2 1 0 1 2 2 2
a t r i c t
2 1 1 2 3 3
h t r i c t
2 2 1 2 3
```

We can reduce computation using diagonal monotonicity:

* Whenever the value on a diagonal \(d\) grows larger than \(k\), we can discard \(d\) from consideration, because we are only interested in values at most \(k\) on the row \(m\).
* We keep track of the smallest undiscarded diagonal \(d\). Each column is computed only up to diagonal \(d+1\).
The position of the smallest undiscarded diagonal on the current column is kept in a variable top.

**Algorithm 3.14:** Ukkonen’s cut-off algorithm

Input: text $T[1..n]$, pattern $P[1..m]$, and integer $k$.

Output: end positions of all approximate occurrences of $P$.

1. top ← $\min(k + 1, n)$
2. for $i \leftarrow 0$ to top do $g_{0,i} \leftarrow i$
3. for $j \leftarrow 1$ to $n$ do $g_{0,j} \leftarrow 0$
4. for $i \leftarrow 1$ to top do
5. for $j \leftarrow 1$ to top do
6. $g_{i,j} \leftarrow \min(g_{i-1,j-1} + 4, 4, g_{i,j-1} + 1, g_{i-1,j} + 1)$
7. while $g_{\text{top},j} > k$ do top ← top - 1
8. if top = m then output $j$
9. else top ← top + 1; $g_{\text{top},j} \leftarrow k + 1$

The time complexity is proportional to the computed area in the matrix $(g_{i,j})$.

- The worst case time complexity is still $O(mn)$ on ordered alphabet.
- The average case time complexity is $O(kn)$. The proof is not trivial.

There are many other algorithms based on diagonal monotonicity. Some of them achieve $O(kn)$ worst case time complexity.

**Example 3.16:** ‘–’ means $-1$, ‘=’ means 0 and ‘+’ means $+1$. 

```
remachin
0 = 0 = 0 = 0 = 0 = 0 = 0 = 0 = 0 = 0
m + + ... = 4 = 4 – 3 – 2 – 1 + 2 + 3 + 4 = 4
h + + + + + = + = + + = + = + + + + + +
5 = 5 = 5 – 4 – 3 – 2 – 1 + 2 + 3 + 4
```

Myers’ Bitparallel Algorithm

Another way to speed up the computation is bitparallelism. Instead of the matrix $(g_{i,j})$, we store $m$ bits.

**Lemma 3.15:** $\Delta h_{ij}, \Delta v_{ij} \in \{-1, 0, 1\}$ for every $i,j$ that they are defined for.

The proof is left as an exercise.

In the standard computation of a cell:

- Input is $g_{i-1,j}$, $g_{i-1,j-1}$, $g_{i,j-1}$ and $\delta(P[i], T[j])$.
- Output is $g_{ij}$.

In the corresponding bitparallel computation:

- **Input:** is $\Delta_{\text{in}} = \Delta h_{i,j-1}$, $\Delta_{\text{h}} = \Delta h_{i,j-1}$ and $\Delta_{\text{out}} = 1 - \delta(P[i], T[j])$.
- **Output:** is $\Delta_{\text{out}} = \Delta h_{i,j}$ and $\Delta_{\text{h}} = \Delta h_{i,j}$.

```
g_{i-1,j-1} \quad \Delta_{\text{in}} \quad g_{i-1,j} \\
\downarrow \quad \Delta_{\text{h}} \quad \uparrow \\
\Delta_{\text{in}} \quad g_{i,j-1} \quad \Delta_{\text{out}} \quad g_{ij}
```

The algorithm does not compute the $\Delta d$ values but they are useful in the proofs.

To enable bitparallel operation, we need two changes:

- The $\Delta v$ and $\Delta a$ values are “trits” not bits. We encode each of them with two bits as follows:

  \[
  \begin{align*}
  P_v &= \begin{cases} 1 & \text{if } \Delta v = +1 \\ 0 & \text{otherwise} \end{cases} \\
  M_v &= \begin{cases} 1 & \text{if } \Delta v = -1 \\ 0 & \text{otherwise} \end{cases} \\
  P_h &= \begin{cases} 1 & \text{if } \Delta h = +1 \\ 0 & \text{otherwise} \end{cases} \\
  M_h &= \begin{cases} 1 & \text{if } \Delta h = -1 \\ 0 & \text{otherwise} \end{cases}
  \end{align*}
  \]

- We replace arithmetic operations ($+$, $-$, min) with Boolean (logical) operations ($\lor$, $\land$, $\neg$).

The computation rule is defined by the following result.

**Lemma 3.17:** If $\text{Eq} = 1$ or $\Delta v_{\text{in}} = -1$ or $\Delta h = -1$, then $\Delta d = 0$, $\Delta v_{\text{out}} = \Delta v_{\text{h}}$ and $\Delta h_{\text{out}} = \Delta v_{\text{in}}$.

Otherwise $\Delta d = 1$, $\Delta v_{\text{out}} = 1 - \Delta v_{\text{h}}$ and $\Delta h_{\text{out}} = 1 - \Delta v_{\text{in}}$.

**Proof:** We can write the recurrence for $g_{ij}$ as

$$g_{ij} = \min\{g_{i-1,j-1} + \delta(P[i], T[j]), g_{i-1,j} + 1, g_{i,j-1} + 1\}$$

Then $\Delta d = g_{ij} - g_{i-1,j} = \min\{1 - \text{Eq}, \Delta v_{\text{out}} + 1, \Delta v_{\text{h}} + 1\}$ which is $0$ if $\text{Eq} = 1$ or $\Delta v = -1$ or $\Delta h = -1$ and $1$ otherwise.

Clearly $\Delta d = \Delta d_{\text{in}} + \Delta d_{\text{out}} = \Delta h_{\text{out}} + \Delta h_{\text{out}}$.

Thus $\Delta d_{\text{in}} = \Delta d - \Delta h_{\text{out}}$ and $\Delta d_{\text{out}} = \Delta d - \Delta v_{\text{out}}$. \qed

Now the computation rules can be expressed as follows.

**Lemma 3.18:**

$$\begin{align*}
P_v &= \Delta d_{\text{in}} \lor (\text{Ph} \land \text{Mh}) \\
M_v &= \Delta d_{\text{out}} \lor (\text{Ph} \land \text{Mh}) \\
P_h &= \Delta d_{\text{in}} \lor (\text{Ph} \land \text{Mh}) \\
M_h &= \Delta d_{\text{out}} \lor (\text{Ph} \land \text{Mh})
\end{align*}$$

**Proof:** We show the claim for $P_v$ and $M_v$ only. $P_h$ and $M_h$ are symmetrical.

By **Lemma 3.17,**

$$\begin{align*}
P_v &= (\neg \text{Eq} \lor \text{Ph} \land \text{Mh}) \lor (\text{Eq} \land \neg \text{Ph} \land \text{Mh}) \\
M_v &= (\neg \text{Eq} \lor \text{Ph} \land \text{Mh}) \lor (\text{Eq} \land \neg \text{Ph} \land \text{Mh}) \\
\text{Ph} &= (\text{Ph} \land \text{Mh}) \lor \text{Xv} \\
\text{Mh} &= (\text{Ph} \land \text{Mh}) \lor \text{Xv}
\end{align*}$$

Because $\Delta d = (\neg \text{Eq} \lor \text{Ph} \land \text{Mh}) = \neg \text{Xv} \land \neg \text{Mh}$,

$$\begin{align*}
P_v &= (\text{Ph} \land \text{Mh}) \lor (\text{Ph} \land \text{Mh}) = (\text{Ph} \land \text{Mh}) = \text{Xv} \lor \neg \text{Mh} \\
M_v &= (\text{Ph} \land \text{Mh}) \lor (\text{Ph} \land \text{Mh}) = (\text{Ph} \land \text{Mh}) = \text{Xv} \lor \neg \text{Mh}
\end{align*}$$

All the steps above use just basic laws of Boolean algebra except the last step, where we use the fact that $\text{Mh}$ and $\text{Ph}$ cannot be 1 simultaneously. \qed
According to Lemma 3.18, the bit representation of the matrix can be computed as follows.

\[
\text{for } i \leftarrow 1 \text{ to } m \text{ do }
\begin{align*}
P_{v0} &\leftarrow 1; M_{v0} \leftarrow 0 \\
&\text{for } j \leftarrow 1 \text{ to } m \text{ do } \\
P_{v0} &\leftarrow 0; M_{v0} \leftarrow 0 \\
&\text{for } i \leftarrow 1 \text{ to } m \text{ do }
\begin{align*}
X_{vij} &\leftarrow Eq_{ij} \lor M_{vij-1} \\
M_{vij} &\leftarrow P_{vij-1} \lor X_{vij} \\
\end{align*}
\text{for } i \leftarrow 1 \text{ to } m \text{ do }
\begin{align*}
X_{vij} &\leftarrow Eq_{ij} \lor M_{vij-1} \\
P_{vij} &\leftarrow M_{hi-1,j} \lor \neg (X_{vij} \lor P_{hi-1,j}) \\
M_{vij} &\leftarrow P_{hi-1,j} \land X_{vij} \\
\end{align*}
\text{This is not yet bitparallel though.}
\]

Lemma 3.19: \(X_{hij} \in \exists i \in [1..m] : Eq_{ij} \land (\forall x \in \{0, 1\} : P_{vij-1})\).

Proof. We use induction on \(i\).

Basis \(i = 1\): The right-hand side reduces to \(Eq_{1j}\), because \(\ell = 1\). By Lemma 3.18, \(X_{hij} = Eq_{ij} \lor M_{hij}\), which is \(Eq_{ij}\) because \(M_{hij} = 0\) for all \(j\).

Induction step: The induction assumption is that \(X_{hij-1}\) is as claimed. Now we have

\[
\exists i \in [1..m] : Eq_{ij} \land (\forall x \in \{0, 1\} : P_{vij-1})
\]

\[
= Eq_{ij} \lor \exists i \in [1..m] : Eq_{ij} \land (\forall x \in \{0, 1\} : P_{vij-1})
\]

\[
= Eq_{ij} \lor P_{vij-1,j} \land \exists i \in [1..m] : Eq_{ij} \land (\forall x \in \{0, 1\} : P_{vij-1,j})
\]

\[
= Eq_{ij} \lor (P_{vij-1,j} \land X_{vij})
\]

(Lemma 3.18) \(= X_{hij}\) (Lemma 3.18)

Lemma 3.20: Denote \(X = X_{hij}, E = Eq_{ij}, P = P_{vij-1}\) and let \(Y = ((E \land P) \lor P) \lor E\). Then \(X = Y\).

Proof. By Lemma 3.19, \(X[i] = 1\) iff and only if

\(a) E[i] = 1\) or
\(b) \exists i \in [1..m] : E[e..i] = 00...01 \land P[e..i-1] = 11...11\).

and \(X[i] = 0\) iff and only if
\(c) E[e..i] = 00...00\) or
\(d) \exists i \in [1..m] : E[e..i] = 00...01 \land P[e..i-1] \neq 11...11\).

We prove that \(Y[i] = X[i]\) in all of these cases:

a) The definition of \(Y\) ends with \(\lor E\) which ensures that \(Y[i] = 1\) in this case.

As a final detail, we compute the bottom row values \(g_{mij}\) using the equalities \(g_{mij} = m \lor g_{m, j-1} + \Delta h_{mij}\).

Algorithm 3.21: Myers' bitparallel algorithm

Input: text \(T[1..n]\), pattern \(P[1..|P|]\), and integer \(k\)

Output: end positions of all approximate occurrences of \(P\)

\(1)\) for \(e \in \Sigma\) do \(B[e] \leftarrow 0\)

\(2)\) for \(i \leftarrow 1 \text{ to } m \text{ do } \bar{B}[P[i]] \leftarrow 1\)

\(3)\) \(P_e \leftarrow 1\); \(M_e \leftarrow 0\); \(g \leftarrow m\)

\(4)\) for \(j \leftarrow 1 \text{ to } m \text{ do }
\begin{align*}
E_{eq} &\leftarrow B[P[j]] \\
X &\leftarrow ((Eq_{eq} \lor P_e) \lor P_e) \lor Eq \\
P &\leftarrow M_e \lor \neg (X \lor (Ph \land \neg P_e))
\end{align*}

\(5)\) \(M_e \leftarrow P_e \lor X\)

\(6)\) \(X_{vij} \leftarrow ((Eq_{eq} \lor P_e) \lor P_e) \lor Eq \\
P &\leftarrow M_e \lor \neg (X \lor (Ph \land \neg P_e))
\end{align*}

\(7)\) \(M_e \leftarrow P_e \lor X\)

\(8)\) \(g \leftarrow g + P_m - M_m\)

\(9)\) \(g \leftarrow g + P_m - M_m\)

\(10)\) \(g \leftarrow g + P_m - M_m\)

\(11)\) \(g \leftarrow g + P_m - M_m\)

\(12)\) \(g \leftarrow g + P_m - M_m\)

\(13)\) if \(g \leq k\) then \(output j\)

To obtain a bitparallel algorithm, the columns \(P_{v0j}, M_{v0j}, X_{vij}, Ph_{ij}, Mh_{ij}, X_{hij}\) and \(Eq_{ij}\) are stored in bitvectors.

Now the second inner loop can be replaced with the code

\[
X_{vij} \leftarrow Eq_{ij} \lor M_{vij-1} \\
P_{vij} \leftarrow (M_{hi-1,j} < 1) \lor \neg (X_{vij} \lor P_{hi-1,j}) \\
M_{vij} \leftarrow P_{hi-1,j} \land X_{vij}
\]

A similar attempt with the for first inner loop leads to a problem:

\[
X_{hij} \leftarrow Eq_{ij} \lor (M_{hij} < 1) \\
P_{hij} \leftarrow (M_{hij} < 1) \lor \neg (X_{hij} \land P_{hi-1,j}) \\
M_{hij} \leftarrow P_{hi-1,j} \land X_{hij}
\]

Now the vector \(M_{hij}\) is used in computing \(X_{hij}\) before \(M_{hij}\) itself is computed! Changing the order does not help, because \(X_{hij}\) is needed to compute \(M_{hij}\).

To get out of this dependency loop, we compute \(X_{hij}\) without \(M_{hij}\) using only \(Eq_{ij}\) and \(P_{vij-1}\) which are already available when we compute \(X_{hij}\).

At first sight, we cannot use Lemma 3.19 to compute even a single bit in constant time, let alone a whole vector \(X_{hij}\). However, it can be done, but we need more bit operations:

- Let \(\oplus\) denote the xor-operation: \(0 \oplus 1 = 1 \oplus 0 = 1\).
- A bitvector is interpreted as an integer and we use addition as a bit operation. The carry mechanism in addition plays a key role. For example \(0001 + 0111 = 1000\).

In the following, for a bitvector \(B\), we will write

\[
B = B[1..m] = B[m]B[m-1]...B[1]
\]

The reverse order of the bits reflects the interpretation as an integer.

b) The following calculation shows that \(Y[i] = 1\) in this case:

\[
E[l..i] = E[0..i] = 00...01 \\
P[l..i] = B[1..i] = b1...11
\]

\[
((E \land P) \lor P) \lor E = (E[0..i] \lor P[0..i]) \lor E[0..i] = b1...11
\]

\[
Y = ((E \land P) \lor P) \lor E[i] = b1...11
\]

where \(b\) is the unknown bit \(P[i]\), \(c\) is the possible carry bit coming from the summation of bits \(l\), \(i\), and \(c\) are their negations.

- Because for all bitvectors \(B, 0 \land B = 0\) \(0 = B \lor B = b\), we get \(Y = ((E \land P) \lor P) \lor E[i] = b1...11\).

- Consider the calculation in case \(b\). A key point there is that the carry bit in the summation travels from position \(l\) to \(i\) and produces \(c\) to position \(i\). The difference in this case is that at least one bit \(P[k], k < i\), is zero, which stops the carry at position \(k\). Thus \(((\bar{E} \land P) \lor P)[0..i] = b\) and \(Y[i] = (b \oplus b) \lor 0 = 0\).

On an integer alphabet, when \(m \leq w\):

- Pattern preprocessing time is \(O(m + \sigma)\).
- Search time is \(O(w)\).

When \(m > w\), we can store each bit vector in \(\lfloor m/w \rfloor\) machine words:

- The worst case search time is \(O(n/\lfloor m/w \rfloor)\).
- Using Ukkonen's cut-off heuristic, it is possible reduce the average case search time to \(O(n/k\lfloor m/w \rfloor)\).

There are also algorithms for bitparallel simulation of a nondeterministic automaton that recognizes the approximate occurrences of the pattern.

Example 3.22:

\[
P = \text{pattern, } k = 3
\]
Another way to utilize Lemma 3.15 \((\Delta_{ij}, \Delta_{vij} \in \{-1, 0, 1\})\) is to use precomputed tables to process multiple matrix cells at a time.

- There are at most 3\(^\text{rd}\) different columns. Thus there exists a deterministic automaton with 3\(^\text{rd}\) states and 3\(^\text{rd}\) transitions that can find all approximate occurrences in \(O(n)\) time. However, the space and construction time of the automaton can be too big to be practical.
- There is a super-alphabet algorithm that processes \(O(\log_s n)\) characters at a time and \(O(\log_s^2 n)\) matrix cells at a time using lookup tables of size \(O(n)\). This gives time complexity \(O(mn/\log^2 n)\).
- A practical variant uses smaller lookup tables to compute multiple entries of a column at a time.

The following lemma shows the property used by the Baeza-Yates–Perleberg algorithm and proves that it satisfies the first condition.

**Lemma 3.23:** Let \(P_1 P_2 \ldots P_{k+1} = P\) be a partitioning of the pattern \(P\) into \(k + 1\) nonempty factors. Any string \(S\) with \(ed(P, S) \leq k\) contains \(P\) as a factor for some \(i \in [k+1]\).

**Proof.** Each single symbol edit operation can change at most one of the pattern factors \(P_i\). Thus any set of at most \(k\) edit operations leaves at least one of the factors untouched. \(\square\)

Let us analyze the average case time complexity of the verification phase.

- The best pattern partitioning is as even as possible. Then each pattern factor has length at least \(r = |m/(k + 1)|\).
- The expected number of exact occurrences of a random string of length \(r\) in a random text of length \(n\) is at most \(n/r^s\).
- The expected total verification time is at most 
  \[
  O\left(\frac{m^2(k+1)n}{\sigma^r}\right) \leq O\left(\frac{m^3n}{\sigma^r}\right)
  \]
  This is \(O(n)\) if \(r \geq 3\log_s m\).
- The condition \(r \geq 3\log_s m\) is satisfied when \((k + 1) \leq m/(3\log_s m + 1)\).

**Theorem 3.24:** The average case time complexity of the Baeza-Yates–Perleberg algorithm is \(O(n)\) when \(k \leq m/(3\log_s m + 1) - 1\).

**Baeza-Yates–Perleberg Filtering Algorithm**

A filtering algorithm for approximate string matching searches the text for factors having some property that satisfies the following conditions:

1. Every approximate occurrence of the pattern has this property.
2. Strings having this property are reasonably rare.
3. Text factors having this property can be found quickly.

Each text factor with the property is a potential occurrence, which is then verified for whether it is an actual approximate occurrence.

Filtering algorithms can achieve linear or even sublinear average case time complexity.

The algorithm has two phases:

- **Filtration:** Search the text \(T\) for exact occurrences of the pattern factors \(P_i\). Using the Aho–Corasick algorithm this takes \(O(n)\) time for a constant alphabet.

- **Verification:** An area of length \(O(m)\) surrounding each potential occurrence found in the filtration phase is searched using the standard dynamic programming algorithm in \(O(m^2)\) time.

The worst case time complexity is \(O(m^2 n)\), which can be reduced to \(O(mn)\) by combining any overlapping areas to be searched.

Many variations of the algorithm have been suggested:

- The filtration can be done with a different multiple exact string matching algorithm.
- The verification time can be reduced using a technique called hierarchical verification.
- The pattern can be partitioned into fewer than \(k + 1\) pieces, which are searched allowing a small number of errors.

A lower bound on the average case time complexity is \(\Omega(n(k + \log_s m)/m)\), and there exists a filtering algorithm matching this bound.

**Summary: Approximate String Matching**

We have seen two main types of algorithms for approximate string matching:

- Basic dynamic programming time complexity is \(O(mn)\). The time complexity can be improved to \(O(kn)\) using diagonal monotonicity, and to \(O(n[m/w])\) using bitparallelism.
- Filtering algorithms can improve average case time complexity and are the fastest in practice when \(k\) is not too large.

Similar techniques can be useful for other variants of edit distance but not always straightforwardly.

**4. Suffix Trees and Arrays**

Let \(T = T[0..n]\) be the text. For \(i \in [0..n]\), let \(T_i\) denote the suffix \(T[i..n]\). Furthermore, for any subset \(C \subseteq [0..n]\), we write \(T_C = \{i \mid i \in C\}\). In particular, \(T_{[0..a]}\) is the set of all suffixes of \(T\).

Suffix tree and suffix array are search data structures for the set \(T_{[0..a]}\):

- Suffix tree is a compact trie for \(T_{[0..a]}\).
- Suffix array is an ordered array for \(T_{[0..a]}\).

They support fast exact string matching on \(T\):

- A pattern \(P\) has an occurrence starting at position \(i\) if and only if \(P\) is a prefix of \(T_i\).
- Thus we can find all occurrences of \(P\) by a prefix search in \(T_{[0..a]}\).

A data structure supporting fast string matching is called a text index. There are numerous other applications too, as we will see later.
The set $T_{[0..n]}$ contains $|T_{[0..n]}| = n+1$ strings of total length $|T_{[0..n]}| = \Theta(n^2)$. It is also possible that $\Sigma LCP(T_{[0..n]}) = \Theta(n^2)$, for example, when $T = \sigma^n$ or $T = XX$ for any string $X$.

- A basic trie has $\Theta(n^2)$ nodes for most texts, which is too much.
- A compact trie with $O(n)$ nodes and an ordered array with $n+1$ entries have linear size.
- A compact ternary trie has $O(n)$ nodes too. However, the construction algorithms and some other algorithms we will see are not straightforward to adapt for it.

Even for a compact trie or an ordered array, we need a specialized construction algorithm, because any general construction algorithm would need $\Omega(\Sigma LCP(T_{[0..n]}))$ time.

As with tries, there are many possibilities for implementing the child operation. We again avoid this complication by assuming that $\sigma$ is constant. Then the size of the suffix tree is $O(n)$.

- There are exactly $n+1$ leaves and at most $n$ internal nodes.
- There are at most $2n$ edges. The edge labels are factors of the text and can be represented by pointers to the text.

Given the suffix tree of $T$, all occurrences of $P$ in $T$ can be found in time $O(|P|+occ)$, where $occ$ is the number of occurrences.

Now we are ready to describe the construction algorithm.

**Algorithm 4.2: Brute force suffix tree construction**

Input: text $T_{[0..n]}$ ($T[n] = \$$)

Output: suffix tree of $T$.

1. create new node root; depth(root) ← 0
2. u ← root; d ← 0 // (u,d) is the active locus
3. for i ← 0 to n do // insert suffix $T_i$
   4. while d = depth(u) and child(u, $[i+d]$) ≠ ∅ do
   5. u ← child(u, $[i+d]$); d ← d + 1
   6. while d < depth(u) and $T[depth(u)] = T[i+d]$ do d ← d + 1
   7. if d < depth(u) then // (u,d) is in the middle of an edge
      8. u ← CreateNode(u,d)
   9. CreateLeaf(u,u)
10. u ← root; d ← 0

CreateLeaf(u,w) // Create leaf representing suffix $T_w$
1. create new leaf w
2. start(w) ← i; depth(w) ← n - i + 1
3. child(u, $[i+d]$) ← w; parent(w) ← u // Set u as parent
4. return w
Suffix links are well defined for all nodes except the root.

**Lemma 4.4:** If the suffix tree of $T$ has a node $u$ representing $T[i..j]$ for any $0 \leq i < j \leq n$, then it has a node $v$ representing $T[i+1..j]$. 

**Proof.** If $u$ is the leaf representing the suffix $T_i$, then $v$ is the leaf representing the suffix $T_{i+1}$.

If $u$ is an internal node, then it has two child edges with labels starting with different symbols, say $a$ and $b$, which means that $T[i..j]a$ and $T[i..j]b$ are both factors of $T$. Then, $T[i+1..j]a$ and $T[i+1..j]b$ are factors of $T$ too, and thus there must be a branching node $v$ representing $T[i+1..j]$. □

Usually, suffix links are needed only for internal nodes. For root, we define $slink(root) = root$.

Theorem 4.6: Let $T$ be a string of length $n$ over an alphabet of constant size. McCreight’s algorithm computes the suffix tree of $T$ in $O(n)$.

Since $d' + 1$ can never be larger than $n$, the total time in the loop on lines (3)-(4) in ComputeSlink is $O(n)$.

### McCreight’s Algorithm

McCreight’s suffix tree construction is a simple modification of the brute force algorithm that computes the suffix links during the construction and uses them as short cuts:

- Consider the situation, where we have just added a leaf $v$, representing the suffix $T_i$, as a child to a node $u_i$.
- The brute force algorithm finds $u_{i+1}$ by traversing from the root. McCreight’s algorithm takes a short cut to $slink(u_i)$.

![Diagram](image)

This is safe because $slink(u_i)$ represents a prefix of $T_{i+1}$.

**Theorem 4.6:** Let $T$ be a string of length $n$ over an alphabet of constant size. McCreight’s algorithm computes the suffix tree of $T$ in $O(n)$ time.

**Proof.** Insertion of a suffix $T_i$ takes constant time except in two points:

- The while loops on lines (4)-(6) traverse from the node $slink(u_i)$ to $u_{i+1}$. Every round in these loops increments $d$. The only place where $d$ decreases is on line (11) and even then by one. Since $d$ can never exceed $n$, the total time on lines (4)-(6) is $O(n)$.
- The while loop on lines (3)-(4) during a call to ComputeSlink($u_i$) traverses from the node $slink(parent(u_i))$ to $slink(u_i)$. Let $d_i$ be the depth of $parent(u_i)$. Clearly, $d_{i+1} \geq d_i + 1$, and every round in the while loop during ComputeSlink($u_i$) increases $d_{i+1}$. Since $d_i$ can never be larger than $n$, the total time in the loop on lines (3)-(4) in ComputeSlink is $O(n)$.

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Applications of Suffix Tree
Let us have a glimpse of the numerous applications of suffix trees.

Exact String Matching
As already mentioned earlier, given the suffix tree of the text, all occurrences of a pattern $P$ can be found in time $O(|P| + occ)$. Even if we take into account the time for constructing the suffix tree, this is asymptotically as fast as Knuth–Morris–Pratt for a single pattern and Aho–Corasick for multiple patterns.

However, the primary use of suffix trees is in indexed string matching, where we can afford to spend a lot of time in preprocessing the text, but must then answer queries very quickly.

Text Statistics
Suffix tree is useful for computing all kinds of statistics on the text. For example:

- Every locus in the suffix tree represents a factor of the text and, vice versa, every factor is represented by some locus. Thus the number of distinct factors in the text is exactly the number of distinct locuses, which can be computed by a traversal of the suffix tree in $O(n)$ time even though the resulting value is typically $\Theta(n^2)$.

- The longest repeating factor of the text is the longest string that occurs at least twice in the text. It is represented by the deepest internal node in the suffix tree.

AC Automaton for the Set of Suffixes
As already mentioned, a suffix tree with suffix links is essentially an Aho–Corasick automaton for the set of all suffixes.

- We saw that it is possible to follow suffix link / failure transition from any locus, not just from suffix tree nodes.

- Following such an implicit suffix link may take more than a constant time, but the total time during the scanning of a string with the automaton is linear in the length of the string. This can be shown with a similar argument as in the construction algorithm.

Thus suffix tree is asymptotically as fast to operate as the AC automaton, but needs much less space.

LCA Preprocessing
The lowest common ancestor (LCA) of two nodes $u$ and $v$ is the deepest node that is an ancestor of both $u$ and $v$. Any tree can be preprocessed in linear time so that the LCA of any two nodes can be computed in constant time. The details are omitted here.

- Let $w_i$ and $w_j$ be the leaves of the suffix tree of $T$ that represent the suffixes $T_i$ and $T_j$. The lowest common ancestor of $w_i$ and $w_j$, $v$, represents the longest common prefix of $T_i$ and $T_j$. Thus $\text{lcp}(T_i, T_j) = \text{depth}(\text{LCA}(w_i, w_j))$, which can be computed in constant time using the suffix tree with LCA preprocessing.

- The longest common prefix of two suffixes $S$ and $T$ from two different strings $S$ and $T$ is called the longest common extension. Using the generalized suffix tree with LCA preprocessing, the longest common extension for any pair of suffixes can be computed in constant time.

Some $O(kn)$ worst case time approximate string matching algorithms use longest common extension data structures.

Approximate String Matching
Several approximate string matching algorithms achieving $O(kn)$ worst case time complexity are based on suffix trees.

Filtering algorithms that reduce approximate string matching to exact string matching such as partitioning the pattern into $k + 1$ factors, can use suffix trees in the filtering phase.

Another approach is to generate all strings in the $k$-neighborhood of the pattern, i.e., all strings within edit distance $k$ from the pattern and search for them in the suffix tree.

The best practical algorithms for indexed approximate string matching are hybrids of the last two approaches. For example, partition the pattern into $\ell \leq k + 1$ factors and find approximate occurrences of the factors with edit distance $[k/\ell]$ using the neighborhood method in the filtering phase.

Generalized Suffix Tree
A generalized suffix tree of two strings $S$ and $T$ is the suffix tree of the string $S \ell T$. It can be computed in linear time.

The generalized suffix tree can also be defined for more than two strings.

Matching Statistics
The matching statistics of a string $S[0..n]$ with respect to a string $T$ is an array $MS[0..n]$, where $MS[i]$ is a pair $(l, p)$ such that

1. $S[l..i]$ is the longest prefix of $S$ that is a factor of $T$, and
2. $T[p..i] = S[i..i + t]$, where $t = \text{start}(i)$.

Matching statistics can be computed by using the suffix tree of $T$ as an AC-automaton and scanning $S$ with it.

- If before reading $S[j]$ we are at the locus $(v, d)$ in the automaton, then $S[j..j + d] = T[j..j + d]$, where $j = \text{start}(i)$. If reading $S[i]$ causes a failure transition, then $MS[i - d] = (d, j)$.

- Following the failure transition decrements $d$ and thus increments $i - d$ by one. Following a normal transition/edge, increments both $i$ and $d$ by one, and thus $i - d$ stays the same. Thus all entries are computed.

From the matching statistics, we can easily compute the longest common factor of $S$ and $T$. Because we need the suffix tree only for $T$, this saves space compared to a generalized suffix tree.

Matching statistics are also used in some approximate string matching algorithms.

Longest Palindrome
A palindrome is a string that is its own reverse. For example, saippuakauppias is a palindrome.

We can use the LCA preprocessed generalized suffix tree of a string $T$ and its reverse $T^R$ to find the longest palindrome in $T$ in linear time.

- Let $kl$ be the length of the longest common extension of $T_{i+1}$ and $T^R_{n-j}$, which can be computed in constant time. Then $T[i - k_{i-1} + k_{i+1}]$ is the longest odd length palindrome with the middle at $i$.

- We can find the longest odd length palindrome by computing $k_i$ for all $i \in [0..n]$ in $O(n)$ time.

- The longest even length palindrome can be found similarly in $O(n)$ time. The longest palindrome overall is the longer of the two.
Suffix Array
The suffix array of a text $T$ is a lexicographically ordered array of the set $T[0..n]$ of all suffixes of $T$. More precisely, the suffix array is an array $SA[0..n]$ of integers containing a permutation of the set $[0..n]$ such that $T[0][SA[0]] < T[0][SA[1]] < \cdots < T[0][SA[n]]$.

A related array is the inverse suffix array $SA^{-1}$ which is the inverse permutation, i.e., $SA^{-1}[i] = j$ for all $i \in [0..n]$. The value $SA^{-1}[i]$ is the lexicographical rank of the suffix $T[i]$. As with suffix trees, it is common to add the end symbol $T[n] = \$$. It has no effect on the suffix array assuming $\$ $ is smaller than any other symbol.

Example 4.7: The suffix array and the inverse suffix array of the text $T = \text{banana}$. | $i$ | $SA[i]$ | $T[SA[i]]$ | $j$ | $SA^{-1}[j]$ |
<table>
<thead>
<tr>
<th></th>
<th></th>
<th></th>
<th></th>
<th></th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>$$</td>
<td>$$</td>
<td>0</td>
<td>$$</td>
</tr>
<tr>
<td>1</td>
<td>5</td>
<td>$\text{a}$</td>
<td>1</td>
<td>3</td>
</tr>
<tr>
<td>2</td>
<td>3</td>
<td>$\text{ana}$</td>
<td>2</td>
<td>6</td>
</tr>
<tr>
<td>3</td>
<td>1</td>
<td>$\text{anaa}$</td>
<td>3</td>
<td>2</td>
</tr>
<tr>
<td>4</td>
<td>0</td>
<td>$\text{banana}$</td>
<td>4</td>
<td>5</td>
</tr>
<tr>
<td>5</td>
<td>4</td>
<td>$\text{na}$</td>
<td>5</td>
<td>1</td>
</tr>
<tr>
<td>6</td>
<td>2</td>
<td>$\text{nana}$</td>
<td>6</td>
<td>0</td>
</tr>
</tbody>
</table>

Exact String Matching

As with suffix trees, exact string matching in $T$ can be performed by a **prefix search** on the suffix array $SA$. The answer can be conveniently given as a **contiguous interval** $[SA[\ell],SA[\ell+1))$ that contains the suffixes with the given prefix.

The interval can be found using string binary search

- If we have the additional arrays $LLCP$ and $RLCP$, the result interval can be computed in $O(l + \log n)$ time.
- Without the additional arrays, we have the same time complexity on average but the worst case time complexity is $O((l \log n)^2)$
- We can then count the number of occurrences in $O(1)$ time, list all occ occurrences in $O(occ)$ time, or list a sample of $k$ occurrences in $O(k)$ time.

An alternative algorithm for computing the interval $SA[\ell,\ell+1)$ is called backward search. It is commonly used with compressed representations of suffix arrays and will be covered in the course Data Compression Techniques.

LCP Array

Efficient string binary search uses the arrays $LLCP$ and $RLCP$. However, for many applications, the suffix array is augmented with the lcp array of Definition 1.11 (Exercise 4).

Example 4.8: The LCP array for $T = \text{banana}$. | $i$ | $SA[i]$ | $LCP[i]$ | $T[SA[i]]$ |
<table>
<thead>
<tr>
<th></th>
<th></th>
<th></th>
<th></th>
</tr>
</thead>
<tbody>
<tr>
<td>0</td>
<td>$$</td>
<td>$$</td>
<td>$$</td>
</tr>
<tr>
<td>1</td>
<td>5</td>
<td>$\text{a}$</td>
<td>1</td>
</tr>
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<td>2</td>
<td>3</td>
<td>$\text{ana}$</td>
<td>2</td>
</tr>
<tr>
<td>3</td>
<td>1</td>
<td>$\text{anaa}$</td>
<td>3</td>
</tr>
<tr>
<td>4</td>
<td>0</td>
<td>$\text{banana}$</td>
<td>4</td>
</tr>
<tr>
<td>5</td>
<td>4</td>
<td>$\text{na}$</td>
<td>5</td>
</tr>
<tr>
<td>6</td>
<td>2</td>
<td>$\text{nana}$</td>
<td>6</td>
</tr>
</tbody>
</table>

LCP Array Construction

The LCP array is easy to compute in linear time using the suffix array $SA$ and its inverse $SA^{-1}$. The idea is to compute the lcp values by comparing the suffixes, but skip a prefix based on a known lower bound for the lcp value obtained using the following result.

Lemma 4.9: For any $i \in [0..n)$, $LCP[SA^{-1}[i-1]] \geq LCP[SA^{-1}[i]] - 1$.

Proof. For each $i \in [0..n)$, let $\Phi(i) = SA^{-1}[i-1] - 1$. Then $T[\Phi(i)]$ is the immediate lexicographical predecessor of $T[i]$ and $LCP[SA^{-1}[i]] = LCP(T[\Phi(i)]).$

- Let $\ell = LCP[SA^{-1}[i-1]]$ and $\ell' = LCP[SA^{-1}[i]]$. We want to show that $\ell' \geq \ell - 1$. If $\ell = 0$, the claim is trivially true.
- If $\ell > 0$, then for some symbol $c$, $T[i-1] = cT[\Phi(i)] = cT[\Phi(i)-1]+1$. Thus $T[\Phi(i)-1] = T[i-1] \cdot cT[\Phi(i)-1] = T[i-1]$.
- If $\Phi(i) = \Phi(i-1) + 1$, then $\ell' = LCP(T[\Phi(i)],T[\Phi(i)+1]) = LCP(T[i-1],T[i]) = \ell - 1$.
- If $\Phi(i) \neq \Phi(i-1) + 1$, then $T[\Phi(i)-1] < T[i]$ and $\ell' = LCP(T[\Phi(i)],T[\Phi(i)+1]) = \ell - 1$.

RMQ Preprocessing

The range minimum query (RMQ) asks for the smallest value in a given range in an array. Any array can be preprocessed in linear time so that RMQ for any range can be answered in constant time.

In the LCP array, RMQ can be used for computing the lcp of any two suffixes.

Lemma 4.11: The length of the longest common prefix of two suffixes $T[i..n]$ and $T[j..n]$ is $\min\{LCP[i], LCP[j], LCP[i]+LCP[j]-1\}$.

The lemma can be seen as a generalization of Lemma 1.31 (Exercise 3) and holds for any sorted array of strings. The proof is left as an exercise.

- The RMQ preprocessing of the LCP array supports the same kind of applications as the LCA preprocessing of the suffix tree, but RMQ preprocessing is simpler than LCA preprocessing.
- The RMQ preprocessed LCP array can also replace the LLCP and RLCP arrays in binary searching.
We will next describe the RMQ data structure for an arbitrary array \(L[1..n]\) of integers.

- We precompute and store the minimum values for the following collection of ranges:
  - Divide \(L[1..n]\) into blocks of size \(\log n\).
  - For all \(0 < \ell < \log(n/\log n)\), include all ranges that consist of \(2^\ell\) blocks. There are \(O(\log n - \frac{\log n}{\log \log n}) = O(n)\) such ranges.
  - Include all prefixes and suffixes of blocks. There are a total of \(O(n)\) of them.

- Now any range \([i..j]\) that overlaps or touches a block boundary can be exactly covered by at most four ranges in the collection.

The minimum value in \([i..j]\) is the minimum of the minimums of the covering ranges and can be computed in constant time.

Enhanced Suffix Array

The enhanced suffix array adds two more arrays to the suffix and LCP arrays to make the data structure fully equivalent to suffix tree.

- The idea is to represent a suffix tree node \(v\) representing a factor \(S_v\) by the suffix array interval of the suffixes that begin with \(S_v\). That interval contains exactly the suffixes that are in the subtree rooted at \(v\).

- The additional arrays support navigation in the suffix tree using this representation: one array along the regular edges, the other along suffix links.

With all the additional arrays the suffix array is not very space efficient data structure any more. Nowadays suffix arrays and trees are often replaced with compressed text indexes that provide the same functionality in much smaller space. These will be covered in the course Data Compression Techniques.

Here are some of the key properties of the BWT.

- The BWT is easy to compute using the suffix array:
  \[ L[i] = \left\{ \begin{array}{ll} \$ & \text{if } SA[i] = 0 \\ T[SA[i] - 1] & \text{otherwise} \end{array} \right. \]

- The BWT is invertible, i.e., \( T \) can be reconstructed from the BWT \( L \) alone. The inverse BWT can be computed in the same time it takes to sort the characters.

- The BWT \( L \) is typically easier to compress than the text \( T \). Many text compression algorithms are based on compressing the BWT.

- The BWT supports backward searching, a different technique for indexed exact string matching. This is used in many compressed text indexes.

BWT will be covered in more detail in the course Data Compression Techniques.

Prefix Doubling

Our first specialized suffix array construction algorithm is a conceptually simple algorithm achieving \(O(n \log n)\) time.

Let \( T^\ell[i] \) denote the text factor \( T[i..\min(i + \ell, n + 1)] \) and call it an \( \ell \)-factor. In other words:

- \( T^\ell[i] \) is the factor starting at \( i \) and of length \( \ell \) except when the factor is cut short by the end of the text.
- \( T^\ell[i] \) is the prefix of the suffix \( T_i \) of length \( \ell \), or \( T_i \) when \( |T_i| < \ell \).

The idea is to sort the sets \( T^\ell[0..n] \) for ever increasing values of \( \ell \).

- First sort \( T^\ell[0..n] \), which is equivalent to sorting individual characters. This can be done in \( O(n \log n) \) time.
- Then, for \( \ell = 1, 2, 4, 8, \ldots \), use the sorted set \( T^\ell[0..n] \) to sort the set \( T^{2\ell}[0..n] \) in \( O(n) \) time.
- After \( O(\log n) \) rounds, \( \ell > n \) and \( T^\ell[0..n] = T[0..n] \), so we have sorted the set of all suffixes.

Ranges \([i..j]\) that are completely inside one block are handled differently.

- Let \( NSV(i) = \min\{k > i \mid L[k] < L[i]\} \) (NSV=Next Smaller Value). Then the position of the minimum value in the range \([i..j]\) is the last position in the sequence \(i, NSV(i), NSV(NSV(i)), \ldots\) that is in the range. We call these the NSV positions for \( i \).
- For each \( i \), store the NSV positions for \( i \) up to the end of the block containing \( i \) as a bit vector \( B(i) \). Each bit corresponds to a position within the block and one if it is an NSV position. The size of \( B(i) \) is \( \log n \) bits and we can assume that it fits in a single machine word. Thus we need \( O(n) \) words to store \( B(i) \) for all \( i \).
- The position of the minimum in \([i..j]\) is found as follows:
  - Turn all bits in \( B(i) \) after position \( j \) into zeros. This can be done in constant time using bitwise shift -operations.
  - The right-most 1-bit indicates the position of the minimum. It can be found in constant time using a lookup table of size \( O(n) \).

All the data structures can be constructed in \( O(n) \) time (exercise).

Burrows–Wheeler Transform

The Burrows–Wheeler transform (BWT) is an important technique for text compression, text indexing, and their combination compressed text indexing.

Let \( T[0..n] \) be the text with \( T[n] = \$ \). For any \( i \in [0..n] \), \( T[i..n]T[0..i] \) is a rotation of \( T \). Let \( M \) be the matrix, where the rows are all the rotations of \( T \) in lexicographical order. All columns of \( M \) are permutations of \( T \).

- The first column \( F \) contains the text characters in order.
- The last column \( L \) is the BWT of \( T \).

**Example 4.12**: The BWT of \( T = \text{banana} \) is \( L = \text{anna}\$\text{ba}\text{a} \).

\[
\begin{array}{cccc}
F & a & n & a & n & a \\
L & b & a & n & a & n & a \\
\end{array}
\]

Suffix Array Construction

Suffix array construction means simply sorting the set of all suffixes.

- Using standard sorting or string sorting the time complexity is \( \Omega(\Sigma LCP(T[0..n])) \).
- Another possibility is to first construct the suffix tree and then traverse it from left to right to collect the suffixes in lexicographical order. The time complexity is \( O(n) \) on a constant alphabet.

Specialized suffix array construction algorithms are a better option, though.

We still need to specify, how to use the order for the set \( T[0..n] \) to sort the set \( T^{\ell}[0..n] \). The key idea is assigning order preserving names (lexicographical names) for the factors in \( T^{\ell}[0..n] \). For \( i \in [0..n] \), let \( N^\ell_i \) be an integer in the range \([0..n]\) such that, for all \( i, j \in [0..n] \):

\[
N^\ell_i \leq N^\ell_j \quad \text{if and only if} \quad T^{\ell}_i \preceq T^{\ell}_j.
\]

Then, for \( \ell > n \), \( N^\ell = SA^{-1}[\ell] \).

For smaller values of \( \ell \), there are many ways of satisfying the conditions and any one of them will do. A simple choice is \( N^\ell = |\{j \in [0..n] \mid T^{\ell}_j \prec T^{\ell}_i\}| \).

**Example 4.13**: Prefix doubling for \( T = \text{banana} \).

<table>
<thead>
<tr>
<th>(N^1)</th>
<th>(N^2)</th>
<th>(N^3)</th>
<th>(N^4)</th>
</tr>
</thead>
<tbody>
<tr>
<td>a</td>
<td>a</td>
<td>a</td>
<td>a</td>
</tr>
<tr>
<td>b</td>
<td>b</td>
<td>b</td>
<td>b</td>
</tr>
<tr>
<td>a</td>
<td>a</td>
<td>a</td>
<td>a</td>
</tr>
<tr>
<td>aa</td>
<td>aa</td>
<td>aa</td>
<td>aa</td>
</tr>
<tr>
<td>b</td>
<td>b</td>
<td>b</td>
<td>b</td>
</tr>
<tr>
<td>$</td>
<td>$</td>
<td>$</td>
<td>$</td>
</tr>
</tbody>
</table>

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Now, given $N^\ell$, for the purpose of sorting, we can use
• $N^\ell_i$ to represent $T^\ell_i$
• the pair $(N^\ell_i, N^\ell_{i+\ell})$ to represent $T^{2\ell}_i$.
Thus we can sort $T^{2\ell}_{[0..n]}$ by sorting pairs of integers, which can be done in $O(n \log n)$ time using LSD radix sort.

**Theorem 4.14:** The suffix array of a string $T[0..n]$ can be constructed in $O(n \log n)$ time using prefix doubling.

- The technique of assigning order preserving names to factors whose lengths are powers of two is called the Karp–Miller–Rosenberg naming technique. It was developed for other purposes in the early seventies when suffix arrays did not exist yet.
- The best practical variant is the Larson–Sadakane algorithm, which uses ternary quicksort instead of LSD radix sort for sorting the pairs, but still achieves $O(n \log n)$ total time.

Let us return to the first phase of the prefix doubling algorithm: assigning names $N^\ell_i$ to individual characters. This is done by sorting the characters, which is easily within the time bound $O(n \log n)$, but sometimes we can do it faster:
- On an ordered alphabet, we can use ternary quicksort for time complexity $O(n \log \sigma_T)$ where $\sigma_T$ is the number of distinct symbols in $T$.
- On an integer alphabet of size $n^\ell$ for any constant $c$, we can use LSD radix sort with radix $n$ for time complexity $O(n)$.

After this, we can replace each character $T[i]$ with $N^\ell_i$ to obtain a new string $T'$.
- The characters of $T'$ are integers in the range $[0..n]$.
- The character $T'[n]$ is 0, the unique, smallest symbol, i.e., $\$.
- The suffix arrays of $T$ and $T'$ are exactly the same.

Thus we can construct the suffix array using $T'$ as the text instead of $T$. As we will see next, the suffix array of $T'$ can be constructed in linear time. Then sorting the characters of $T$ to obtain $T'$ is the asymptotically most expensive operation in the suffix array construction of $T$ for any alphabet.

**Assume that**
- $|C| \leq an$ for a constant $a < 1$, and
- excluding the recursive call, all steps in the algorithm take linear time.

Then the total time complexity can be expressed as the recurrence $t(n) = O(n) + t(\alpha n)$, whose solution is $t(n) = O(n)$.

To make the scheme work, the set $C$ must satisfy two nontrivial conditions:
1. There exists an appropriate reduced string $R$.
2. Given sorted $T_C$, the suffix array of $T$ is easy to construct.

Finding sets $C$ that satisfy both conditions is difficult, but there are two different methods leading to two different algorithms:
- DC3 uses difference cover sampling
- SAIS uses induced sorting

**Difference Cover Sampling**

A difference cover $D_i$ modulo $q$ is a subset of $[0..q]$ such that all values in $[0..q]$ can be expressed as a difference of two elements in $D_i$ modulo $q$. In other words:
\[ [0..q] = \{i - j \mod q \mid i, j \in D_i \}. \]

**Example 4.15:** $D_1 = \{1, 2, 4\}$
\[
\begin{align*}
1 - 1 &= 0 \\
2 - 1 &= 1 \\
4 - 2 &= 2 \\
4 - 1 &= 3 \\
\end{align*}
\]
In general, we want the smallest possible difference cover for a given $q$.

- For any $q$, there exist a difference cover $D_i$ of size $O(\sqrt{q})$.
- The DC3 algorithm uses the simplest non-trivial difference cover $D_3 = \{1, 2\}$.

**Algorithm 4.18: DC3**

**Step 0:** Choose $C$.
- Use difference cover $D_3 = \{1, 2\}$.
- For $k \in \{0, 1, 2\}$, define $C_k = \{i \in [0..n] \mid i \mod 3 = k\}$. Let $C = C_0 \cup C_1 \cup C_2$.

**Example 4.19:**
\[
\begin{align*}
T[i] &\quad \text{y a b b a d a b b a d o}\$ \\
C_0 &\quad = \{0, 3, 6, 9, 12\}, C_1 = \{1, 4, 7, 10\}, \quad C_2 = \{2, 5, 8, 11\} \\
&\quad \text{and } C = \{1, 2, 4, 5, 7, 8, 10, 11, 12\}.
\end{align*}
\]

A difference cover sample $T_C$ of suffixes, where $C = \{i \in [0..n] \mid (i \mod q) \in D_i\}$.

**Example 4.16:** If $T = \text{banana}$ and $D_3 = \{1, 2\}$, then $C = \{1, 2, 4, 5\}$ and $T_C = \{\text{a}, \text{ana}, \text{anaa}, \text{anaa}, \text{a}\}$.

Once we have sorted the difference cover sample $T_C$, we can compare any two suffixes in $O(q)$ time. To compare suffixes $T_i$ and $T_j$:
- If $i \in C$ and $j \in C$, then we already know their order from $T_C$.
- Otherwise, find $i$ such that $i + \ell \in C$ and $j + \ell \in C$. There always exists such $\ell \in [0..q)$. Then compare $T_i = T_{i+\ell}$ and $T_j = T_{j+\ell}$.
  - That is, compare first $T[1]..T[\ell]$ to $T[1]..T[\ell]$, and if they are the same, then $T_{i+\ell}$ to $T_{j+\ell}$ using the sorted $T_C$.

**Example 4.17:** $D_3 = \{1, 2\}$ and $C = \{1, 2, 4, 5, \ldots\}$
\[
\begin{align*}
T_0 &= T[0]T_1 = T[7]T[1]T_2 = T[0]T_3 & T_4 = T[0]T_1 \\
\end{align*}
\]

**Step 1:** Sort $T_C$.
- For $k \in \{1, 2\}$, construct the strings $R_k = (T^1_1, T^1_2, T^1_3, \ldots, T^1_{\max^k})$ whose characters are 3-factors of the text, and let $R = R_1R_2$.
  - Replace each factor $T^k_i$ in $R$ with an order preserving name $N^k_i \in [1..|R|]$.
  - The names can be computed by sorting the factors with LSD radix sort in $O(n)$ time. Let $R'$ be the result appended with 0.
  - Construct the inverse suffix array $SA_{R'}^{-1}$ of $R'$. This is done recursively using DC3 unless all symbols in $R'$ are unique, in which case $SA_{R'}^{-1} = R'$.
  - From $SA_{R'}^{-1}$, we get order preserving names for suffixes in $T_C$.
  - For $i \in C$, let $N_i = SA_{R'}^{-1}[i]$, where $j$ is the position of $T^k_j$ in $R$.
  - For $i \notin C$, let $N_i = \$. Also let $N_{\max^k} = N_{\max} = \$.

**Example 4.20:**
\[
\begin{align*}
R &\quad \text{aba bda bab abd ad}\$ \\
&\quad 1 2 4 7 4 6 3 8 0 \\
SA_{R'}^{-1} &\quad 1 2 5 7 4 6 3 8 0 \\
&\quad i 0 1 2 3 4 5 6 7 8 9 10 11 12 13 14 \\
T[i] &\quad \text{y a b b a d a b b a d o}\$ \\
N_i &\quad 1 4 2 6 5 3 7 8 1 0 0
\end{align*}
\]
Step 2(a): Sort $T^{-C}$.
- For each $i \in C$, we represent $T_i$ with the pair $(T_i, N_{i+1})$. Then $T_i \leq T_j \iff (T_i, N_{i+1}) \leq (T_j, N_{j+1})$.
- Note that $N_{i+1} \neq \perp$ for all $i \in C$.
- The pairs $(T_i, N_{i+1})$ are sorted by LSD radix sort in $O(n)$ time.

Example 4.21:

\[
\begin{array}{cccccccccccc}
  i & 0 & 1 & 2 & 3 & 4 & 5 & 6 & 7 & 8 & 9 & 10 \\
\end{array}
\]

$T_12 \prec T_6 \prec T_9 \prec T_3 \prec T_0$ because $(\$ , 0) < (a , 5) < (a , 7) < (b , 2) < (y , 1)$.

Step 2(b): Merge $T_i$ and $T_C$.
- Use comparison based merging algorithm needing $O(n)$ comparisons.
- To compare $T_i \in T_i$ and $T_C \in T_C$, we have two cases:
  - $i \in C_1$: $T_i \leq T_C \iff (\{i\}, N_{i+1}) \leq (\{i\}, N_{i+1})$
  - $i \in C_2$: $T_i \leq T_C \iff (T_i, T_i + 1, N_{i+1}) \leq (T_C, T_C + 1, N_{i+1})$
- Note that none of the $N$-values is $\perp$.

Example 4.22:

\[
\begin{array}{cccccccccccc}
  i & 0 & 1 & 2 & 3 & 4 & 5 & 6 & 7 & 8 & 9 & 10 \\
\end{array}
\]

$T_1 \prec T_3 \prec T_5 \prec T_7$ because $(a , 4) < (a , 5)$.

$T_1 \prec T_5$ because $(b , 6) < (b , 7)$.

Induced Sorting

Define three types of suffixes via, +, and * as follows:

- $C^-$ = \{ $i \in [0..n]$ | $T_i > T_{i+1}$ \}
- $C^+$ = \{ $i \in [0..n]$ | $T_i < T_{i+1}$ \}
- $C^* = \{ i \in C^+ | i-1 \in C^- \}$

Example 4.24:

\[
\begin{array}{cccccccccccc}
  i & 0 & 1 & 2 & 3 & 4 & 5 & 6 & 7 & 8 & 9 & 10 \\
\end{array}
\]

The basic idea of induced sorting is to use information about the order of $T_i$ to induce the order of the suffix $T_{i-1} = T_{i-1}[..T_i]$. The main steps are:

1. Sort the sets $C_a^-, a \in [1..\sigma]$.
2. Use $C^+ a$, $a \in [1..\sigma]$, to induce the order of the sets $C^+ a$, $a \in [1..\sigma]$. The suffixes involved in the induction steps can be indentified using the following rules (proof is left as an exercise).

Lemma 4.25: For all $a \in [1..\sigma]$:

(a) $i-1 \in C_a^+$ iff $i > 0$ and $T[i -1] = a$ and one of the following holds
- $i = n$
- $i \in C^+$
- $i \in C^-$ and $T[i-1] > T[i]$

(b) $i-1 \in C_a^+$ iff $i > 0$ and $T[i -1] = a$ and one of the following holds
- $i \in C^-$ and $T[i-1] < T[i]$
- $i \in C^+$ and $T[i-1] < T[i]$.

Inducing +--type suffixes goes similarly but in reverse order so that again $i$ is always inserted before $i-1$:

1. Set $C_a^+$ empty for all $a \in [1..\sigma]$.
2. For all suffixes $T_i$ such that $i-1 \in C^+$ in descending lexicographical order, append $i-1$ into $C_{T[i-1]}^-$.

By Lemma 4.25(a), Step 2 can be done by checking the relevant conditions for all $i \in C_a^+ C_a^+ C_a^+ C_a^+ C_a^-$.

Algorithm 4.26: InduceMinusSuffix

Input: Lexicographically sorted lists $C_a^-, a \in [1..\sigma]$
Output: Lexicographically sorted lists $C_a^+, a \in [1..\sigma]$

(1) for $a \in [1..\sigma]$ do $C_a^- \leftarrow \emptyset$
(2) for $a \leftarrow a - 1$ downto 1 do
  (3) for $i \in C_a^+$ in reverse order do // include elements added during loop
    (4) if $i > 0$ and $T[i -1] \leq a$ then pushfront($i-1, C_{T[i-1]}^-$)
  (5) if $i > 0$ and $T[i-1] < a$ then pushfront($i-1, C_{T[i-1]}^-$)
(6) for $i \in C_a^+$ do pushback($i-1, C_{T[i-1]}^-$)

Note that since $T_{i-1} > T_i$ by definition of $C^-$, we always have $i$ inserted before $i-1$.

We still need to explain how to sort the +--type suffixes. Define

$F[i] = \min \{ k \in [0..\sigma] | k \in C^+ \text{ or } k = n \}$
$S_i = T[i..F[i]]$
$S_i' = S_i$

where $\epsilon$ is a special symbol larger than any other symbol.

The main idea is that $S_i' < S_i$.

Algorithm 4.27: InducePlusSuffixes

Input: Lexicographically sorted lists $C_a^-, a \in [1..\sigma]$
Output: Lexicographically sorted lists $C_a^+, a \in [1..\sigma]$

(1) for $a \in [1..\sigma]$ do $C_a^+ \leftarrow \emptyset$
(2) for $a \leftarrow a - 1$ downto 1 do
  (3) for $i \in C_a^+$ in reverse order do // include elements added during loop
    (4) if $i > 0$ and $T[i -1] \leq a$ then pushfront($i-1, C_{T[i-1]}^+$)
  (5) if $i > 0$ and $T[i-1] < a$ then pushfront($i-1, C_{T[i-1]}^+$)
(6) for $i \in C_a^+$ do pushback($i-1, C_{T[i-1]}^+$)
Algorithm 4.29: SAIS

Step 0: Choose $C$.
- Compute the types of suffixes. This can be done in $O(n)$ time based on Lemma 4.25.
- Set $C = \bigcup_{i \in [1, \sigma]} C_i \cup \{n\}$. Note that $|C| \leq n/2$, since for all $i \in C$, $i - 1 \notin C^* \subseteq C$.

Example 4.30:

\begin{align*}
T[i] & \quad 0 \quad 1 \quad 2 \quad 3 \quad 4 \quad 5 \quad 6 \quad 7 \quad 8 \quad 9 \quad 10 \quad 11 \quad 12 \quad 13 \quad 14 \\
\text{type of } T_i & \quad - \quad - \quad - \quad * \quad - \quad - \quad - \quad - \quad * \quad + \quad - \quad - \quad - \quad - \\
C^*_i & = \{2, 5, 8\}, C^*_2 = C^*_p = C^*_s = \emptyset, C = \{2, 5, 8, 14\}.
\end{align*}

Step 1: Sort $T_C$.
- Sort the strings $S'_i, i \in C^*$. Since the total length of the strings $S'_i$ is $O(n)$, the sorting can be done in $O(n)$ time using LSD radix sort.
- Assign order preserving names $N_i \in [1 : |C| - 1]$ to the string $S'_i$ so that $N_i \leq N_j$ iff $S'_i \leq S'_j$.
- Construct the sequence $R = N_i N_i \ldots N_i 0$, where $i_1 < i_2 < \cdots < i_k$ are the $*$-type positions.
- Construct the suffix array $SA_R$ of $R$. This is done recursively unless all symbols in $R$ are unique, in which case a simple counting sort is sufficient.

The order of the suffixes of $R$ corresponds to the order of $*$-type suffixes of $T$. Thus we can construct the lexicographically ordered lists $C^*_a, a \in [1, \sigma]$.

Example 4.31:

\begin{align*}
T[i] & \quad 0 \quad 1 \quad 2 \quad 3 \quad 4 \quad 5 \quad 6 \quad 7 \quad 8 \quad 9 \quad 10 \quad 11 \quad 12 \quad 13 \quad 14 \\
N_i & \quad 2 \quad 2 \quad 1 \quad 1 \quad 1 \quad 1 \quad 0 \quad 0 \quad 0 \quad 0 \quad 0 \quad 0 \quad 0 \quad 0 \\
R & = \text{issi} \sigma \text{issi} \sigma \text{ippii} \sigma \sigma = 2210, SA_R = (3, 2, 1, 0), C^*_1 = (8, 5, 2)
\end{align*}

Summary: Suffix Trees and Arrays

The most important data structures for string processing:
- Designed for indexed exact string matching.
- Used in efficient solutions to a huge variety of different problems.

Construction algorithms are among the most important algorithms for string processing:
- Linear time for constant and integer alphabet.

Often augmented with additional data structures:
- suffix links, LCA preprocessing
- LCP array, RMQ preprocessing, BWT, ...

Theorem 4.33: Algorithm SAIS constructs the suffix array of a string $T[0..n]$ in $O(n)$ time plus the time needed to sort the characters of $T$.

In Step 1, to sort the strings $S'_i, i \in C^*$, SAIS does not actually use LSD radix sort but the following procedure:
1. Construct the sets $C^*_a, a \in [1, \sigma]$ in arbitrary order.
2. Run InduceMinusSuffixes to construct the lists $C^*_a, a \in [1, \sigma]$.
3. Run InducePlusSuffixes to construct the lists $C^*_a, a \in [1, \sigma]$.
4. Remove non-*$a$-type positions from $C^*_1 C^*_2 \ldots C^*_a$.

With this change, most of the work is done in the induction procedures. This is very fast in practice, because all the lists $C^*_a$ are accessed sequentially during the procedures.

The currently fastest suffix sorting implementation in practice is probably divsufsort by Yuta Mori. It sorts the $*$-type suffixes non-recursively in $O(n \log n)$ time and then continues as SAIS.